AUTOMATA THEORY AND COMPUTABILITY (Effective from the academic year 2018 -2019) SEMESTER – V			
Course Code	18CS54	**CIE Marks**	40
Number of Contact Hours/Week	3:0:0	**SEE Marks**	60
Total Number of Contact Hours	40	**Exam Hours**	03
CREDITS –3			

Course Learning Objectives: This course (18CS54) will enable students to:

- Introduce core concepts in Automata and Theory of Computation
- Identify different Formal language Classes and their Relationships
- Design Grammars and Recognizers for different formal languages
- Prove or disprove theorems in automata theory using their properties
- Determine the decidability and intractability of Computational problems

Module 1	**Contact Hours**
Why study the Theory of Computation, Languages and Strings: Strings, Languages. A Language Hierarchy, Computation, **Finite State Machines (FSM)**: Deterministic FSM, Regular languages, Designing FSM, Nondeterministic FSMs, From FSMs to Operational Systems, Simulators for FSMs, Minimizing FSMs, Canonical form of Regular languages, Finite State Transducers, Bidirectional Transducers. **Textbook 1: Ch 1,2, 3,4, 5.1 to 5.10** **RBT: L1, L2**	08
Module 2	
Regular Expressions (RE): what is a RE?, Kleene"s theorem, Applications of REs, Manipulating and Simplifying REs. Regular Grammars: Definition, Regular Grammars and Regular languages. Regular Languages (RL) and Non-regular Languages: How many RLs, To show that a language is regular, Closure properties of RLs, to show some languages are not RLs. **Textbook 1: Ch 6, 7, 8: 6.1 to 6.4, 7.1, 7.2, 8.1 to 8.4** **RBT: L1, L2, L3**	08
Module 3	
Context-Free Grammars(CFG): Introduction to Rewrite Systems and Grammars, CFGs and languages, designing CFGs, simplifying CFGs, proving that a Grammar is correct, Derivation and Parse trees, Ambiguity, Normal Forms. Pushdown Automata (PDA): Definition of non-deterministic PDA, Deterministic and Non-deterministic PDAs, Non- determinism and Halting, alternative equivalent definitions of a PDA, alternatives that are notequivalent to PDA. **Textbook 1: Ch 11, 12: 11.1 to 11.8, 12.1, 12.2, 12,4, 12.5, 12.6** **RBT: L1, L2, L3**	08
Module 4	
Algorithms and Decision Procedures for CFLs: Decidable questions, Un-decidable questions. **Turing Machine**: Turing machine model, Representation, Language acceptability by TM, design of TM, Techniques for TM construction. Variants of Turing Machines (TM), The model of Linear Bounded automata. **Textbook 1: Ch 14: 14.1, 14.2, Textbook 2: Ch 9.1 to 9.8** **RBT: L1, L2, L3**	08
Module 5	
Decidability: Definition of an algorithm, decidability, decidable languages, Undecidable languages, halting problem of TM, Post correspondence problem. Complexity: Growth rate	08

of functions, the classes of P and NP, Quantum Computation: quantum computers, Church-Turing thesis. **Applications:** G.1 Defining syntax of programming language, Appendix J: Security **Textbook 2: 10.1 to 10.7, 12.1, 12.2, 12.8, 12.8.1, 12.8.2** **Textbook 1: Appendix: G.1(only), J.1 & J.2** **RBT: L1, L2, L3**	

Course Outcomes: The student will be able to :

- Acquire fundamental understanding of the core concepts in automata theory and Theory of Computation
- Learn how to translate between different models of Computation (e.g., Deterministic and Non-deterministic and Software models).
- Design Grammars and Automata (recognizers) for different language classes and become knowledgeable about restricted models of Computation (Regular, Context Free) and their relative powers.
- Develop skills in formal reasoning and reduction of a problem to a formal model, with an emphasis on semantic precision and conciseness.
- Classify a problem with respect to different models of Computation.

Question Paper Pattern:

- The question paper will have ten questions.
- Each full Question consisting of 20 marks
- There will be 2 full questions (with a maximum of four sub questions) from each module.
- Each full question will have sub questions covering all the topics under a module.
- The students will have to answer 5 full questions, selecting one full question from each module.

Textbooks:

1. Elaine Rich, Automata, Computability and Complexity, 1st Edition, Pearson education,2012/2013
2. K L P Mishra, N Chandrasekaran , 3rd Edition, Theory of Computer Science, PhI, 2012.

Reference Books:

1. John E Hopcroft, Rajeev Motwani, Jeffery D Ullman, Introduction to AutomataTheory, Languages, and Computation, 3rd Edition, Pearson Education, 2013
2. Michael Sipser : Introduction to the Theory of Computation, 3rd edition, Cengage learning,2013
3. John C Martin, Introduction to Languages and The Theory of Computation, 3rd Edition, Tata McGraw –Hill Publishing Company Limited, 2013
4. Peter Linz, "An Introduction to Formal Languages and Automata", 3rd Edition, Narosa Publishers, 1998
5. Basavaraj S. Anami, Karibasappa K G, Formal Languages and Automata theory, Wiley India, 2012
6. C K Nagpal, Formal Languages and Automata Theory, Oxford University press, 2012.

Faculty can utilize open source tools (like JFLAP) to make teaching and learning more interactive.

MODULE 1

FINITE AUTOMATA

Finite automata are basic modules of a computer system, which addresses simple real time problems. Figure 1.1 shows conceptual view of finite automata. It has an input tape, read head and finite controller. Input tape is divided into finite cells and each cell stores one input symbols. Read head is used to read the input symbol from the input tape. The finite controller used to record the next state of finite automat.

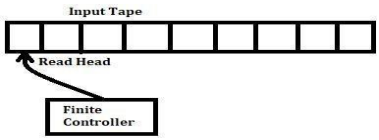


Figure 1.1: Conceptual view of finite automata

Before discussing finite automata in detail, it is very much important to know the basic requirement for the finite automata. They are

i. Sets.
ii. Graphs.
iii. Languages.
iv. Star closure.
v. Plus closure.

Sets (S):

A set is collection of elements, each element has its own identity. A set can be represented by enclosing its elements in curly braces. for example, the set of small letters a, b, c is shown as

$S= \{a, b, c\}$

Set {2, 4, 6,…} denotes the set of all positive even integers. we can use more explicit notation, in which we write

$S=\{i \mid i \text{ is } > 0 \text{ and is even}\}$

We can perform following operations on sets, they are

a. union (U),
b. intersection (∩), and
c. difference (−)
d. complementation defined as

$S_1 \cup S_2 = \{x \mid x \in S_1 \text{ or } x \in S_2\}$

$S_1 \cap S_2 = \{x \mid x \in S_1 \text{ and } x \in S_2\}$

$S_1 - S_2 = \{x \mid x \in S_1 \text{ and } x \notin S_2\}$

The complement of a set S, denoted by $\dot{S}$ consists of all elements not in S. To make this meaningful, we need to know what the universal set U of all possible elements is. If U is specified, then

$\dot{S}=\{x \mid x \in U, x \notin S\}$

The set with no elements, called the empty set or the null set, is denoted by ε. From the definition of a set, it is obvious that

S U ε = S - ε =S

S ∩ ε = ε

ε̃= U

De Morgan's laws: The following useful identities, known as De Morgan's laws.

$\overline{S_1 \cup S_2} = S_1 \cap S_2$

$\overline{S_1 \cap S_2} = S_1 \cup S_2$

are needed on several occasions.

Subset and Proper Subset: A set S_1 is said to be a subset of S, every element of S_1 is also an element of S. We write this as $S_1 \subseteq S$. If $S_1 \subseteq S$, but S contains an element not in S_1, we say that S_1 is a proper subset of S; we write this as $S_1 \subset S$.

Disjoint Sets: If S_1 and S_2 have no common element, that is, $S_1 \cap S_2 = \varepsilon$ then the sets are said to be disjoint.

Finite and Infinite sets: A set is said to be finite if it contains a finite number of elements; otherwise it is infinite. The size of a finite set is the number of elements in it; this is denoted by |S|.

Power- Set: A given set has many subsets. The set of all subsets is called the power-set and is denoted by 2^s. Observe that 2^s is a set of sets.

If S is the set {a, b, c}, then its powerset is

2^s= {ε,{a}, {b}, {c}, {a, b},{b, c},{a, c}, {a, b. c}}

Here |S| = 3 and $|2^s|$ = 8.

Cartesian product: Ordered sequences of elements from other sets are said to be the Cartesian product. For the Cartesian product of two sets, which itself is a set of ordered pairs, we write

$S = S_1 \times S_2$

Let S1 = {2, 4} and S2 = {2, 3, 5, 6}. Then S1 × S2 = {(2, 2), (2, 3), (2, 5), (2, 6), (4, 2), (4, 3), (4, 5), (4, 6)}. Note that the order in which the elements of a pair are written matters. The pair (4, 2) is in S1 × S2, but (2, 4) is not.

Graphs:

A graph is a construct consisting of two finite sets, the set V = {υ1, υ2,…, υn} of vertices and the set E = {e1, e2,…, em} of edges. Each edge is a pair of vertices from V, for instance,

$e_i=(v, v_j)$.

Languages:

A finite, nonempty set of symbols are called the alphabet and is represented with symbol ∑. From the individual symbols we construct strings, which are finite sequences of symbols from the alphabet. For example, if the alphabet ∑ = {0}, then ∑* to denote the set of strings obtained by concatenating zero or more symbols from ∑.

We can write ∑*= {ε, 0, 00, 000 …}

If $\sum = \{0, 1\}$
0 and 1 are input alphabets or symbols, then:
($0^2=00$, $0^3=000$, $1^2=11$, $1^3=111$, where as in mathematics they are numbers $0^2=0$, $1^2=1$)
$L=\sum^*=\{ \varepsilon, 0, 1, 00, 01, 10, 11, 000, 001, 010.....\}$

Kleen closure(*closure):
*closure = $\{\sum^0 U \sum^1 U \sum^2 U \sum^3 \cdots\cdots\}$
$\sum^0=\{ \varepsilon\}$ $\sum^1=\{0, 1\}$ $\sum^2=\{00, 01, 10, 11\}$
$\sum^3=\{000, 001, 010, 011, 100, 101, 110, 111\}$
Therefore $\sum^*=\{ \varepsilon , 0, 1, 00, 01, 10, 11, 000, 001, 010, 011, 100, 101,.....\}$
$^+$Closure on $\sum = \{0, 1\}$
$\sum^+ = \{\sum^1 U \sum^2 U \sum^3 ...\}$
$\sum^* = \sum^+ + \varepsilon$

Therefore $\sum^+ = \sum^* - \varepsilon$

Formal Languages and Automata Theory: IT is a mathematical model used to design computer programs and sequential logic circuits. It is state machines which captures and makes transition to accept or reject. Finite automata has input tape, read head and finite controller **Input tape**: It is divided into finite cell and each cell is used to store one input symbol. **Read head**: Read head is used to read input symbol from input tape. **Finite Controller**: Finite controller records, the transition to next state. The default value of finite controller is initial state q_0. States are represented with circles; they are of three types,

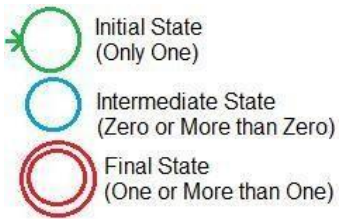


Figure 1.2: Different types of states used in finite automata

States are connected by lines called transition. It has two ends, one end is called initiation point and the other end is called termination point (>).

Initiation ⟶ Termination

The combination of transitions and states is called a transition diagram.

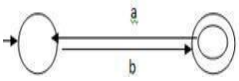


Figure 1.3: A typical transition diagram

Figure 1.4 shows a transition diagram to map cities and their roads.

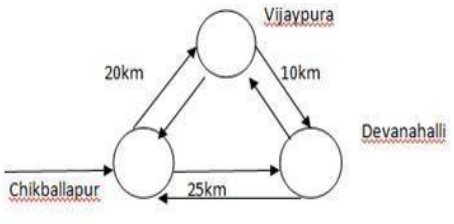


Figure 1.4: Finite automata with real time example.

Finite Automata are classified into two types. They are:
1) Deterministic Finite Automata (DFA)
2) Non Deterministic Finite Automata (NFA)
In Deterministic Finite Automata number of outgoing transitions are well defined from each state depending upon input symbols. For example, if there is one Input symbol, then you find one outgoing transition from each state, for two input symbols two outgoing transitions are present and so on.

Example 1.1: Draw the DFA, the input symbol is ∑= {0}

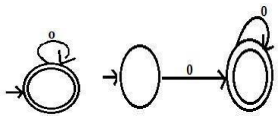


Figure 1.5: DFA's for input symbol ∑= {0}.

Example 1.2: Draw the DFA, the input symbol is ∑= {0, 1}

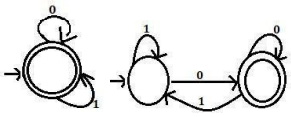


Figure 1.6: DFA's for input symbol ∑= {0, 1}.

Definition of DFA: we define DFA as,

M_{DFA} = (Q, ∑, δ, Initial state, Final state)

Where, Q is set of states, ∑ =set of input symbols, δ: transition function is $Q x \Sigma \rightarrow Q$

From the table 1.

$Q=\{q_0,q_1,q_2\}$ ∑ ={a, b, c}

δ:
δ $(q_0, a)=q_1$ δ $(q_0,b)=q_0$ δ $(q_0,c)=q_2$
δ $(q_1,a)=q_1$ δ $(q_1,b)=q_2$ δ $(q_1,c)=q_0$
δ $(q_2,a)=q_2$ δ $(q_2,b)=q_1$ δ $(q_2,c)=q_2$
Note: Initial State is q_0 and final state is q_2

Table 1.1: Transition table

Q\∑	a	b	c
→ q_0	q_1	q_0	q_2
q_1	q_1	q_2	q_0
*q_2	q_2	q_1	q_2

Nondeterministic Finite Automata (NFA):

If outgoing transitions are not dependent on number of input symbols, then such automata are called as Nondeterministic Finite Automata.

For example: ∑ = {0, 1}

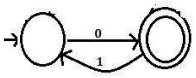


Figure 1.7: Nondeterministic Finite Automata for ∑ = {0, 1}

There is no outgoing transition from initial state, for a symbol '1 'and similarly there is no outgoing transition from final state for '0'.

Transition Function for string:

Note: 1) δ is the transition function, used for symbols.

2) For the string (Set of symbols), $\hat{\delta}$ is used

{ } = ε

{a}=a

$\hat{\delta}$ (q_0, ε)= q_0 //q_0 is reading empty string, string is empty therefore, transition to same state q_0.

We already Proved $\hat{\delta}(q_0,\varepsilon) = q_0$

$\hat{\delta}$ (q_0, a)= $\hat{\delta}$ (q_0, εa)= δ ($\hat{\delta}$ (q_0, ε),a)= δ (,a)

Therefore $\hat{\delta}$ (q_0, a)= δ (q_0, a)

Therefore for single symbol string $\hat{\delta}$ = δ

$\hat{\delta}$ (q_0, wa)

Where 'w' is the multi-symbol string and 'a' is a symbol, hence we separate the transition function for strings and symbols as given below:

δ ($\hat{\delta}$ (q_0, w), a)

Example 1.3: Write a transition table for a DFA shown in figure 1.8.

DFA over alphabet ∑ = {a, b}

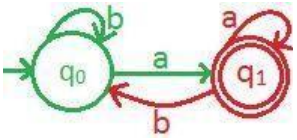


Figure 1.8: Transition diagram o0f a DFA.

Table 1.2: Transition table for the figure 1.8

Q \ ∑	a	b
→q0	δ(q0,a)	δ(q0, b)
*q1	δ(q1, a)	δ(q1,b)

Q \ ∑	a	b
→q0	q1	q0
*q1	q1	q0

M_{DFA} =(Q, ∑, δ , IS, FS)
Q ={ q_0, q_1}
∑ ={a ,b}
δ (q_0, a) =q_1 S δ q_0,b) =q_0
δ (q_1, a) =q_1 δ (q_1, b) =q_0 IS
=q_0 FS={q_0}

Example 1.4: Construct transition table for the diagram shown in figure 1.9, over the alphabet ∑={0, 1, 2}

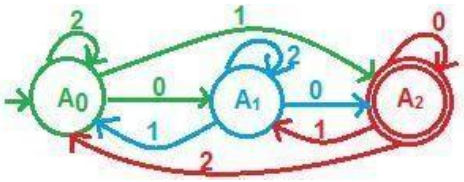


Figure 1.9: Transition diagram

Table 1.3: Transition table for the figure 1.9

A \ ∑	0	1	2
→A0	A1	A2	A0
A1	A2	A0	A1
*A2	A2	A1	A0

We can define above DFA as: M_{DFA}=(A, ∑, δ, IS, {FS})

δ:
δ (A_0, 0)=A_1 δ (A_0, 1)=A_2 δ(A_0, 2)=A_0
δ (A_1, 0)=A_2 δ (A_1, 1)=A_0 δ(A_1, 2)=A_1

$\delta(A_2, 0)=A_2$ $\delta(A_2, 1)=A_1$ $\delta(A_2, 2)=A_0$

$A=\{A_0,A_1,A_2\}$; $\sum=\{0,1,2\}$; $IS=A_0$; $FS=A_2$

String Acceptance:

The first symbol of string begins with initial state. Later, traverses zero or more than zero intermediate state(s) for the intermediate symbol(s). Finally last symbol of the string ends with final state of automata. This is called string accepted or otherwise string is not accepted.

Example 1.5: Find out whether the given string is accepted by the DFA. Assume the string as abab. From the given string $\sum = \{a,b\}$. Note that always string is traversed by left side to right side,

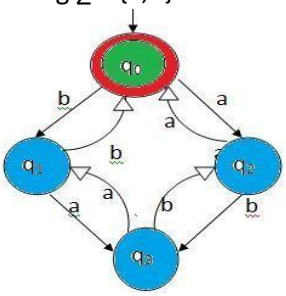


Figure 1.10: Transition diagram for example 1.5.

$\hat{\delta}(q_0, a)= \delta(q_0, a)= q_2$ — 1

$\hat{\delta}(q_0, ab)= \delta(\hat{\delta}(q_0, a), b)$ from 1 $\hat{\delta}(q_0, a)= q_2$

$= \delta(q_2, b) = q_3$ — 2

$\hat{\delta}(q_0, aba)= \delta(\hat{\delta}(q_0, ab), a)$ **from equation 2** $\hat{\delta}(q_0, ab)= q_3$

$= \delta(q_3, a)=q_1$

$\hat{\delta}(q_0, abab)= \delta(\hat{\delta}(q_0, aba), b)$

$= \delta(\hat{\delta}(q_1, b)= q_0$

Given string is accepted by a DFA.

Find out whether given string 10110 is accepted by DFA.

From given string $\sum=\{0,1\}$

10110

$(q_0, 1)=\delta(q_0, 1)=q_2$ ------- ①

10110

$(q_0, 10)=\delta(\delta(q_0, 1),0)=\delta(q_2, 0)=q_2$ ------- ②

10110 $(q_0, 101)=(\delta(q_0, 10),1)=\delta(q_2, 1)=q_1$

10110 $(q_0, 1011)=\delta(\delta(q_0, 101),1)=\delta(q_1, 1)=q_2$

10110 $(q_0, 10110)=\delta(\delta(q_0, 1011),0)=\delta(q_2, 1)=q_2$

q2 is a final state hence given string 10110 is accepted.

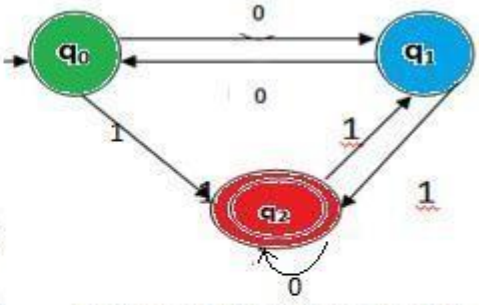


Figure 1.11 : Transition diagram

Example 1.6: Find out the language accepted by the DFA.

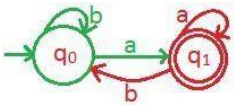


Figure 1.12: Transition diagram of Example 1.6.

$\delta(q_0, a)=q_1$; $\hat{\delta}(q_0, aa)=\delta(\hat{\delta}(q_0, a),a)=\delta(q_1, a)=q_1$

L={a, aa, ba, baa, aba, …..}

In general we can write:

L= {w |w is a string ending with symbol a}

Example 1.7: Find out the language accepted by the DFA.

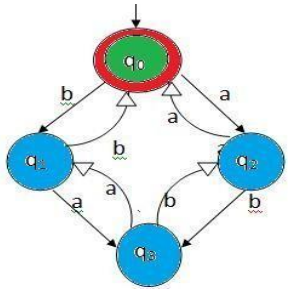


Figure 1.13: Transition diagram of Example 1.7

Initial state is also a final state, hence it accept empty string ε.

$\hat{\delta}(q_0, \varepsilon)=q_0$;

$\hat{\delta}(q_0, aa)=\delta(\hat{\delta}(q_0, a)a)=\delta(q_2, a)=q_0$;

$\hat{\delta}(q_0, bb)=\delta(\hat{\delta}(q_0, b)b)=\delta(q_1, b)=q_0$;

Therefore language L={ ε, aa, bb, aabb, abab…}

In general we can write L={w |w is a string consist of even number of a's and even number of b's}

Design of DFA :

Example 1.8: *Design a DFA which accept all the strings ends with symbol a over the alphabet ∑={a, b}.*

L={a, aa, ba, aba, bba…}

i. Draw the NFA for the smallest string a of the language.
The NFA for the string a is

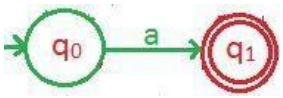


Figure 1.14 : NFA for the string a

Convert all states of NFA to DFA, that is from each state two outgoing transitions, one for symbol 'a' another for 'b'. From initial state only one outgoing transition on symbol 'a', one more outgoing transition is required that is on 'b'. If the condition in the given problem ends with, the outgoing

transition on the other symbol b, definitely loop to a initial state.

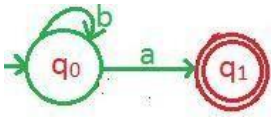


Figure 1.14: second outgoing transition from state q_0

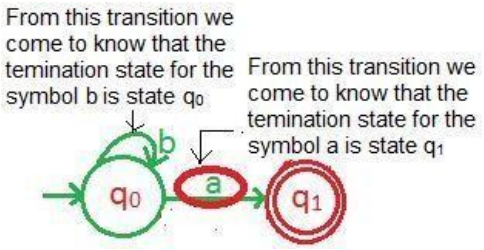
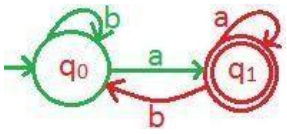


Figure 1.15 : q_0 is termination for symbol b and q_1 is termination for symbol a

Termination state for symbol 'a' is q_1 and for 'b' it is q_0. Outgoing transition from q_1 on symbol 'a' is loop, and for 'b' it is to q_0 .



Figyre 1.16 :DFA which accept all the strings ends with a.

We can define above DFA as M_{DFA} =(Q, ∑, δ , IS, {FS})
Q ={ q_0, q_1}
∑ ={a , b}
δ (q_0, a) =q_1 δ (q_0,b) =q_0 δ
(q_1, a) =q_1 δ (q_1, b) =q_0 IS
=q_0 FS={q_1}

Example 1.8: Design a DFA which accept all the strings ends with substring 01 over the alphabet

∑={0, 1}.

L={01, 001, 101, 0001,1001…}.

i. Draw the NFA for the smallest string 01 of the language.
The NFA for the string 01 is

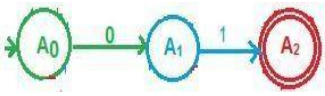


Figure 1.17: NFA for the string 01.

ii. Convert all states of NFA to DFA, that is from each state two outgoing transitions, one for symbol '0' another for '1'. From initial state only one outgoing transition on symbol '0', one

more outgoing transition is required that is on '1'. If the condition in the given problem is, ends with, the outgoing transition on the other symbol 1, definitely loop to a initial state.

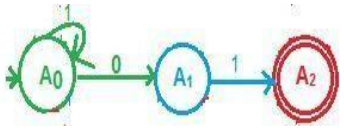


Figure 1.18 : conversion of NFA state A_0 to DFA

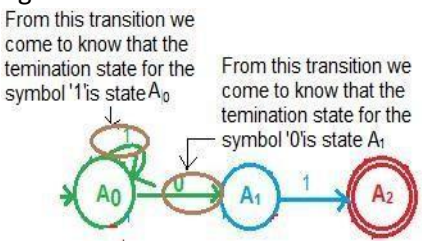
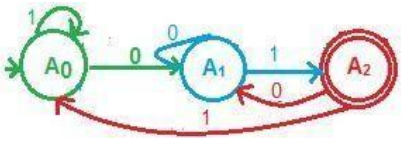


Figure 1.19 : A_0 is termination for symbol 1 and A_1 is termination for symbol 0

Termination state for symbol '0' is A_1 and for '1' it is A_0. Outgoing transition from A_1 on symbol '0' is loop. From A_2 on symbol '0' to A_1 and on symboll '1' to A_0.



Figyre 1.20 :DFA which accept all the strings ends with 01.

We can define above DFA as $M_{DFA}=(A, \sum, \delta, IS, \{FS\})$

$A=\{ A_0, A_1, A_2\}$; $\sum=\{0, 1\}$

δ :

$\delta (A_0, 0)=A_1$ $\delta (A_0, 1)=A_0$ δ

$(A_1, 0)=A_1$ δ $(A_1, 1)=A_2$ δ

$(A_2, 0)=A_1$ δ $(A_2, 1)=A_0$

$IS=A_0$; $FS=A_2$

Example 1.9: Design a DFA which accept all the strings ends with substring 012 over the alphabet

$\sum=\{0, 1, 2\}$.

L={012, 0012, 1012, 00012,10012…}.

i. Draw the NFA for the smallest string 012 of the language.
The NFA for the string 012 is

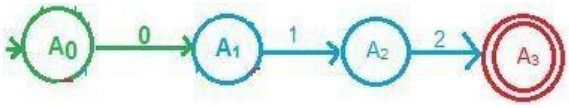


Figure 1.21: NFA for the string 012.

ii. Convert all states of NFA to DFA, that is from each state three outgoing transitions, one for symbol '0' second for '1' third from '2'. From initial state only one outgoing transition on symbol '0', two more outgoing transitions are required that is on '1' and '2'. If the condition

in the given problem ends with, the outgoing transition on the other symbols '1'and '2', definitely loop to a initial state.

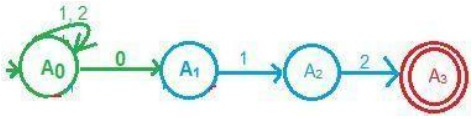


Figure 1.22: conversion of NFA state A_0 to DFA state.

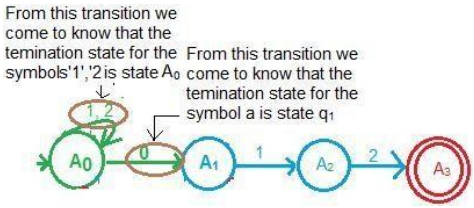


Figure 1.23: A_0 is termination for symbol 1, 2 and A_1 is termination for symbol 0

Termination state for symbol '0' is A_1 and for '1' and '2' is A_0. Outgoing transition from A_1 on symbol '0' is loop and for'2' is A_0. From A_2 on symbol '0' to A_1 and on symbol '1' to A_0. From A_3 on symbol '0' to A_1 and on symbols '1' and '2' to A_0.

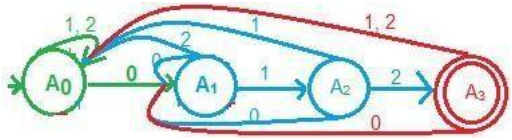


Figure 1.24 :DFA which accept all the strings ends with 012.

We can define above DFA as $M_{DFA}=(A, \sum, \delta, IS, \{FS\})$

$A=\{ A_0, A_1, A_2, A_3\}$; $\sum=\{0, 1, 2\}$

δ :

$\delta (A_0, 0)=A_1$ $\delta (A_0, 1)=A_0$ $\delta(A_0, 2)=A_0$
$\delta (A_1, 0)=A_1$ $\delta (A_1, 1)=A_2$ $\delta(A_1, 2)=A_0$
$\delta (A_2, 0)=A_1$ $\delta (A_2, 1)=A_0$ $\delta(A_2, 2)=A_3$
$\delta (A_3, 0)=A_1$ $\delta (A_3, 1)=A_0$ $\delta(A_3, 2)=A_0$
$IS=A_0$; $FS=A_3$

Example 1.10: Design a DFA which accept all the strings begins with symbol 'a' over the alphabet

$\sum=\{a, b\}$.

L={a, aa, ab, abb, aba, aaa…}

i. Draw the NFA for the smallest string a of the language.
The NFA for the string a is

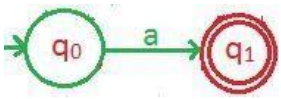


Figure 1.25: NFA for the string a.

Convert all states of NFA to DFA, that is from each state two outgoing transitions, one for symbol 'a' another for 'b'. From initial state only one outgoing transition on symbol 'a', one more outgoing transition is required that is on 'b'. If the condition in the given problem is, begin with, the outgoing transition on the other symbol 'b', is to new state called dead state.

Dead state: In finite automata the string travers from initial state to final state, in some situation we make finite automata to die called dead state. Dead state is a non-final state, from which all outgoing transitions are loop to the same state.

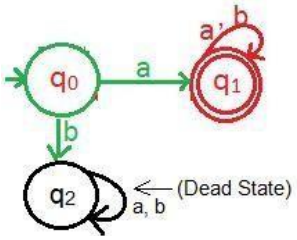


Figure 1.26 : DFA which accept all the string begin with a.

ii. Condition is begin, outgoing transitions from final state is loop.

We can define above DFA as M_{DFA} =(Q, ∑, δ , IS, {FS})

Q ={ q_0, q_1, q_2}

∑ ={a , b}

δ (q_0, a) =q_1 δ (q_0,b) =q_2 δ

(q_1, a) =q_1 δ (q_1, b) =q_1 δ

(q_2, a) =q_2 δ (q_2, b) =q_2 IS

=q_0 FS={q_1}

Example 1.11: Design a DFA which accept all the strings begins with substring 01 over the alphabet

∑={0, 1}.

L={01, 010, 011, 0100, 0101…}.

i. Draw the NFA for the smallest string 01 of the language.
The NFA for the string 01 is

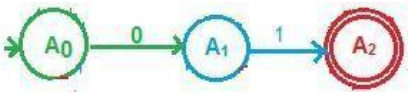


Figure 1.27: NFA for the string 01.

ii. Convert all states of NFA to DFA, that is from each state two outgoing transitions, one for symbol '0' another for '1'. From initial and intermediate states only one outgoing transition, one more outgoing transition is required. If the condition in the given problem is, begin with, the outgoing transition from both the states to new state called dead state.

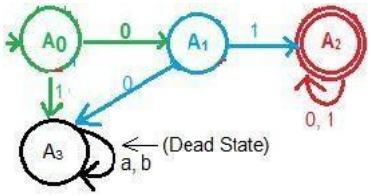


Figure 1.28 : DFA which accept all the string begin with 01.

iii. Condition is begin, outgoing transitions from final state A_2 on symbol '0' and '1' is loop.

We can define above DFA as $M_{DFA}=(A, \sum, \delta, IS, \{FS\})$

$A=\{ A_0, A_1, A_2 A_3\}$; $\sum=\{0, 1\}$

δ :

$\delta (A_0, 0)=A_1$ $\delta (A_0, 1)=A_0$

$\delta (A_1, 0)=A_1$ $\delta (A_1, 1)=A_2$

$\delta (A_2, 0)=A_1$ $\delta (A_2, 1)=A_0$ δ

$(A_3, 0)=A_3$ δ $(A_3, 1)=A_3$

$IS=A_0$; $FS=A_2$

Example 1.12: Design a DFA which accept all the strings begins with 01 or ends with 01 or both over the alphabet $\sum=\{0, 1\}$.

L={01, 010, 011, 0100, 001, 101, 0001,1001, 0101, 01101…}

We have already designed both the DFAs, joining of these DFAs, is the solution for the given problem.

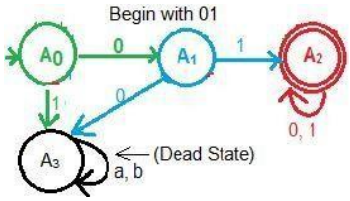


Figure 1.29: DFA which accept all the string begin with 01.

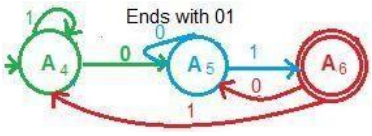


Figure 1.30: DFA which accept all the string ends with 01.

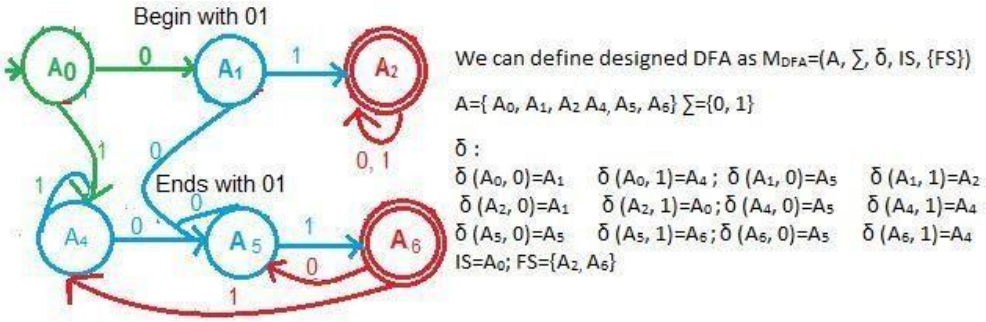


Figure 1.31: DFA which accept all the string begin with 01, end with 01 begin and end with 01.

Example 1.13: Design a DFA which accept all the binary strings divisible by 2.

L={ ε, 10, 100, 110, …}

To design DFA we need to make the following assumptions.

$q_0 \rightarrow 0 \rightarrow q_0$

$q_0 \rightarrow 1 \rightarrow q_1$

$q_0 \rightarrow 10 \rightarrow q_0$

$q_0 \rightarrow 11 \rightarrow q_1$

$q_0 \rightarrow 100 \rightarrow q_0$

$q_0 \rightarrow 101 \rightarrow q_1$

From the above assumption, we can draw the following DFA.

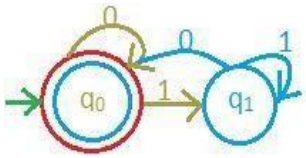


Figure 1.32: DFA which accepts all the strings divisible by 2.

We can define above DFA as M_{DFA} =(Q, ∑, δ , IS, {FS})
Q ={ q_0, q_1}
∑ ={0, 1}
δ (q_0, 0) =q_0 δ (q_0,1) =q_1 δ
(q_1, 0) =q_0 δ (q_1, 1) =q_1 IS
=q_0 FS={q_0}
Example:

Example 1.14: Design a DFA which accept all the binary strings divisible by 5.

L={ ε, 101, 1010, 1111, …}

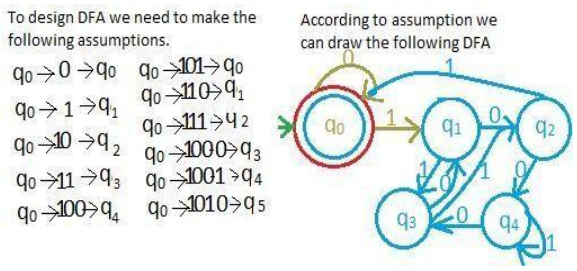


Figure 1.33: DFA which accept all the strings divisible by 5.

We can define, designed DFA as $M_{DFA}=(A, \sum, \delta, IS, \{FS\})$

$A=\{ q_0, q_1, q_2 q_3, q_4\}$ $\sum=\{0, 1\}$

δ :
$\delta (q_0, 0)=q_0$ $\delta (q_0, 1)=q_1$; $\delta (q_1, 0)=q_2$ $\delta (q_1, 1)=q_3$
$\delta (q_2, 0)=q_4$ $\delta (q_2, 1)=q_0$; $\delta (q_3, 0)=q_1$ $\delta (q_3, 1)=q_2$
$\delta (q_4, 0)=q_3$ $\delta (q_4, 1)=q_4$;

$IS=q_0$; $FS=q_0$

Example 1.15: Design a DFA which accept string over x, y every block of length 3 contains at least one x.

To design a DFA we need to make following assumptions,

1. If there is no 'x' in the string, such a string should not be accepted.

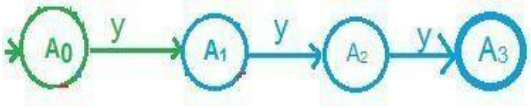


Figure 1.34: NFA which is not accepting three y's string.

2. One 'x' in the string, accepted by finite automata.

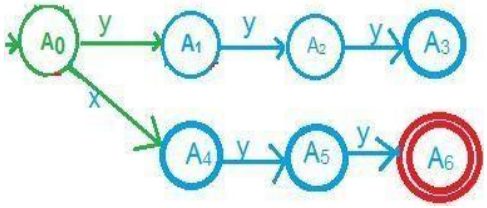


Figure 1.35: NFA which is accepting three symbol string in which one symbol is x.

3. The position of A_1 and A_5 is same as A_0 and A4. In between A_0 and A_4 the symbol is 'x', inbetwwen A_1 and A_5 is also x and so on.

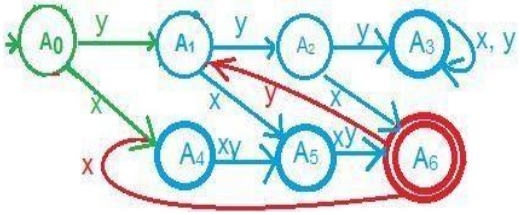


Figure 1.36: *DFA which accept string over x, y if every block of length 3 contains at least one x.*

We can define designed DFA as $M_{DFA}=(A, \sum, \delta, IS, \{FS\})$

$A= \{ A_0, A_1, A_2, A_4, A_4, A_5, A_6\}$ $\sum=\{x, y\}$

δ :
$\delta (A_0, x)= A_4$ $\delta (A_0, y)= A_1$; $\delta (A_1, x)=A_5$ $\delta (A_1, y)=A_2$
$\delta (A_2, x)=A_6$ $\delta (A_2, y)=A_3$; $\delta (A_3, x)=A_3$ $\delta (A_3, y)=A_3$
$\delta (A_4, x)=A_5$ $\delta (A_3, y)=A5$; $\delta (A_5, x)=A_6$ $\delta (A_5, y)=A_6$;

$\delta(A_6, x)=A_4$ $\delta(A_6, y)=A_1$

IS=A_0; FS={A_6}

Example 1.16: Design a DFA which accept even number of r and even number of s over the symbol ∑={r, s}.

i. Draw separate DFA, one accept even number of r's and another one accept even number of s's.

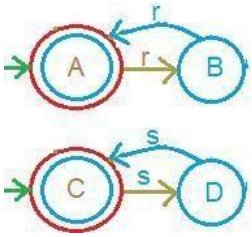


Figure 1.37 : *DFA's accept even symbols string over r, s.*

ii. Join two DFA, we get four new states AB, AC, BC, BD. Designed DFA accept even number of r and even number of s, out of four states AB satisfied this condition, hence this state is a final state.

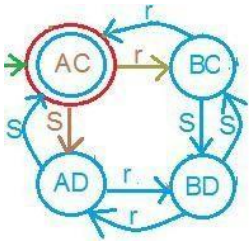


Figure 1.38: *DFA which accept even number of r and s.*

We can define designed DFA as M_{DFA}=(State, ∑, δ, IS, {FS})

State={ AC, AD, BC, BD} ∑={r, s}

δ :
δ (AC, r)= BC δ (AC, s)= AD ; δ (BC, r)=AC δ (BC, s)=BD
δ (AD, r)=BD δ (AD, s)=BD; δ (BD, r)=AD δ (BD, s)=BC

IS=AC; FS={ AC}

Example 1.17: Design a DFA which accept all the strings described by the language L={$N_0(w)\geq 1$ and $N_1(w)=2$}

L={011, 110, 0011, 1010, 11100, 10101...}

i. Exchange I symbol of first string with the last symbol of the first string, the string obtained is the second string of the language.
ii. Draw NFA for two smallest strings, for both NFA initial and final states are same.

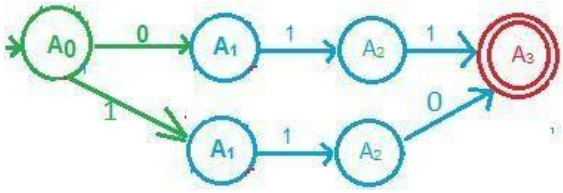


Figure 1.39: Two DFAs for the string 011 and 110 with same initial and final state.

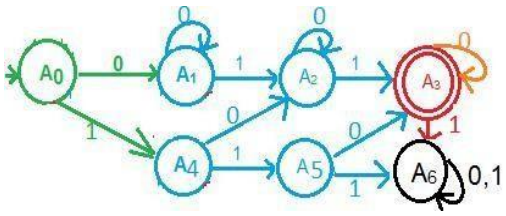


Figure 1.40: DFA which accept all the strings described by the language L={$N_0(w) \geq 1$ and $N_1(w)=2$}.

We can define designed DFA as $M_{DFA}=(A, \sum, \delta, IS, \{FS\})$

A={ A_0, A_1, A_2 A_3, A_4, A_5, A_6} ∑={0, 1}

δ :

δ (A_0, 0)=A_1 δ (A_0, 1)=A_4 ; δ (A_1, 0)=A_0 δ (A_1, 1)=A_2
δ (A_2, 0)=A_2 δ (A_2, 1)=A_3 ; δ (A_4, 0)=A_2 δ (A_4, 1)=A_5
δ (A_5, 0)=A_3 δ (A_5, 1)=A_6 ; δ (A_6, 0)=A_6 δ (A_6, 1)=A_6

IS=A_0; FS={A_3}

Nondeterministic Finite Automata (NFA)

In DFA the outgoing transitions are defined in advance, where as in NFA the outgoing transitions are not defined in advance. In DFA, the string is either accepted or not accepted, where as in NFA both are possible. The finite automata shown in the figure 1.41 is a NFA. There is no outgoing transition for the symbol 'a' from the state q_1.

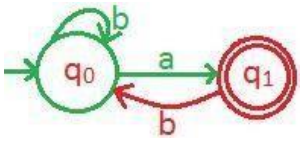


Figure 1.41: Nondeterministic Finite Automata

We can define above NFA as M_{DFA} =(Q, ∑, δ , IS, {FS})
Q ={ q_0, q_1,}
∑ ={a , b}
δ:
$Q^*\sum \rightarrow 2^Q$
δ (q_0, a) =q_1 δ (q_0,b) =q_0
δ (q_1, b) =q_0
δ (q_2, a) =q_2 δ (q_2, b) =q_2
IS =q_0 FS={q_1}

The NFA shown in figure 1.42 is, accepting as well as not accepting the string hod.

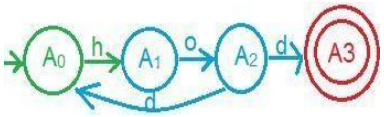


Figure 1.42:Non-deterministic finite automata

Example 1.18: Design NFA which accept all strings, if second symbol from right side is 1 over the alphabet ∑={0, 1}.

L={ 10, 110, 010, 0110…}

i. First step draw the final state.

Figure 1.43: Final state

ii. Draw the intermediate state and transition for the symbols '0' and '1'.

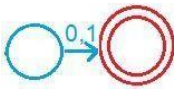


Figure 1.44: Intermediate and final states

iii. Draw the initial state and transition for the symbol '1'.

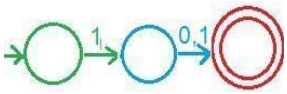


Figure 1.45: Non-deterministic finite automata accepting two strings 10, 11.

iv. Designed NFA has to accept all the strings, if second symbol from right side is symbol '1', to do this, mark loop to initial state on symbols '0' and '1'.

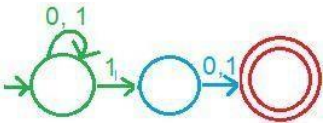


Figure 1.46: NFA which accepts all strings, if second symbol from right side is 1 over the alphabet ∑={0, 1}.

v. Name the states and stop.

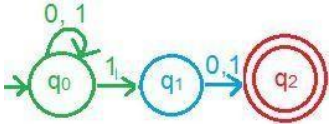


Figure 1.47: NFA which accept all strings, if second symbol from right side is 1 over the alphabet ∑={0, 1}.

Definition of NFA: NFA can be defined as

M_{DFA} =(Q, ∑, δ , IS, {FS})
Q ={ q_0, q_1, q_2}
∑ ={0, 1}
δ:
δ (q_0, 0) =q_0 δ (q_0,1) ={q_0, q_1}
δ (q_1, 0) =q_2
δ (q_1, 1) =q_2
IS =q_0 FS={q_2}

Equivalence of NFA and DFA:

We have already seen both NFA and DFA, both are same in the definition but they differ in transition function. If number of states in NFA has two states q_0 and q_1. Here, q_0 is initial state and q_1 is final state. The equivalent states in DFA's are *q_0], [q_1] and [q_0, q_1]. Where q_0 is initial state and (q_1), (q_0, q_1) are the final states. The pair (q_0, q_1) is final state because; one of its states is final state. They are represented as shown in figure 1.48

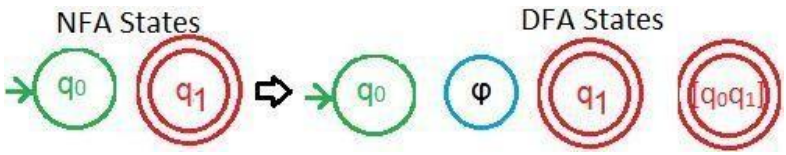


Figure 1.48: Equivalence states of NFA and DFA.

Example 1.19: Convert the NFA shown in figure 1.49 into its equivalent DFA.

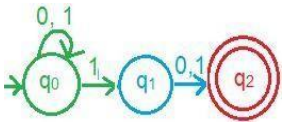


Figure 1.49: Non-deterministic finite automata of Example 1.19.

The conversion begins with the initial state of the NFA which is also an initial state of DFA. Conversion generates number of new states, for every new state we are finding the outgoing transitions fordifferent symbols.

Table 1.4: Different tables for the conversion of NFA to DFA.

Q \ Σ	0	1
→q_0	q_0	[q_0 q_1]

New state

Q \ Σ	0	1
→q_0	q_0	[q_0 q_1]
[q_0q_1]	[q_0q_2]	[$q_0q_1q_2$]

New state

New states

δ([q_0q_1], 0)=δ(q_0, 0)∪δ(q_1, 0)
=q_0 ∪q_2
=q_0q_2
δ([q_0q_1], 1)=δ(q_0, 1)∪δ(q_1,1)
=q_0 q_1∪q_2
=$q_0q_1q_2$

Q \ Σ	0	1
→q0	q0	[q0 q1]
[q0q1]	q0q2	q0q1q2
q0q2	q0	[q0q1]

$\delta([q_0q_2], 0)=\delta(q_0, 0)\cup\delta(q_2, 0)$
$=q_0 \cup \emptyset$
$=q_0$
$\delta([q_0q_2], 1)=\delta(q_0, 1)\cup\delta(q_2, 1)$
$=q_0 q_1 \cup \emptyset$
$=q_0q_1$

Q \ Σ	0	1
→q0	q0	[q0 q1]
[q0q1]	q0q2	q0q1q2
[q0q2]	q0	[q0q1]
[q0q1q2]	[q0q2]	[q0q1q2]

$\delta([q_0q_1q_2],0)=\delta(q_0, 0)\cup\delta(q_1, 0)\cup\delta(q_2, 0)$
$=q_0 \cup q_2 \cup \emptyset$
$=q_0q_2$
$\delta([q_0q_1q_2],1)=\delta(q_0, 1)\cup\delta(q_1, 1)\cup\delta(q_2, 1)$
$=q_0 q_1 \cup q_2 \cup \emptyset$
$=q_0q_1q_2$

No further new states; mark the final state of DFA. Any pair of state which consists of final state of the NFA, the particular pair is a final state in DFA.

Q \ Σ	0	1
→q0	q0	[q0 q1]
[q0q1]	q0q2	q0q1q2
*[q0q2]	q0	[q0q1]
*[q0q1q2]	[q0q2]	[q0q1q2]

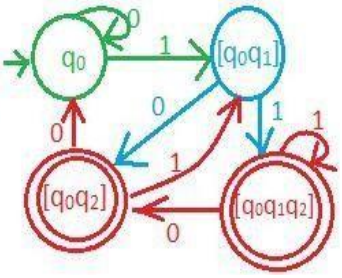


Figure 1.50: Equivalence DFA of figure 1.49

Application of Finite automata:

Finite automat is used to search a text. There are two methods used to search text, they are:

i. Nondeterministic Finite Automata(NFA)
ii. Deterministic Finite Automata(DFA)

To search a text “shine” and “hire” the NFA shown in figure 1.51 is used

The input alphabets are ∑={s, h, i, n, e, r }

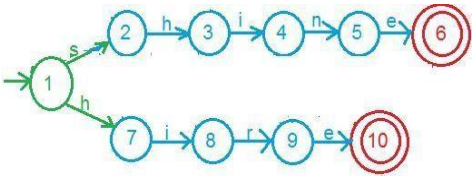


Figure 1.51: NFA used to search strings shine and hire

To search a text “shine” and “hire” the DFA shown in figure 1.52 is used

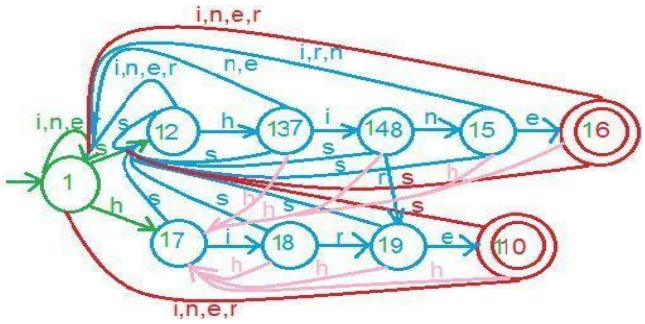


Figure 1.52: DFA used to search strings ending with shine and hire

Example 1.20: Design a finite automata, that reads string made up of letters HONDA, and recognize those string that contain the word HOD as a substring.

i. Find out the alphabet from the string HONDA. The input symbols are ∑={H, O, N, D, A}
ii. Draw the NFA for the substring HOD.

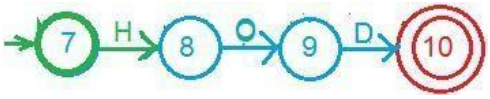


Figure 1.53: NFA used to search string HOD

iii. Convert NFA state of figure 1.53 into DFA states.

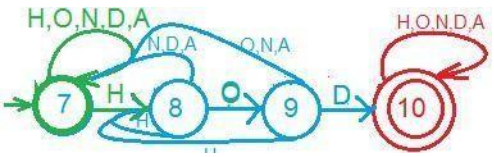


Figure 1.54: DFA accept all the strings ends with HOD over the symbols {H, O, N, D, A}

NFA with ε-transition:

Till now whatever automata we studied were characterized by a transition on the symbols. But there are some finite automata having transition for empty string also, such NFAs are called NFA- ε transition.

ε-closure (q_0): states which are closed / related to state q_0 on empty transition(ε) is called ε-closure (q_0).

In DFA and NFA $\hat{\delta}(q_0, \varepsilon)=q_0$, where as in ε-NFA

i. $\hat{\delta}(q, \varepsilon)$= ε-closure($\hat{\delta}(q, \varepsilon)$)= ε-closure(q).
ii. $\delta(q, x)$= ε-closure(δ($\delta(q, \varepsilon)$, x)= ε-closure(δ(ε-closure(q), x). $\delta \neq \delta$
iii. $\delta(q, wx)$= ε-closure(δ($\delta(q, w)$, x).

Example 1.21: Find out ε-closure of all the states of ε-NFA shown in figure 1.55.

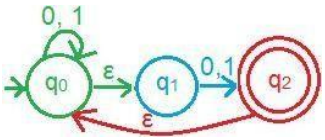


Figure 1.55: NFA with ε-transition of Example 1.21.

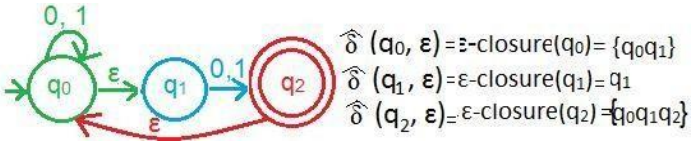


Figure 1.56: ε-closure of all the states of figure 1.55

Example 1.22: Convert the ε-NFA shown in figure 1.57 into NFA

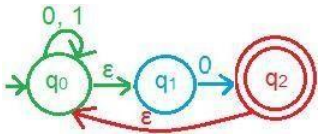


Figure 1.57: ε-NFA of Example 1.22.

First find out the ε-closure of all the states.

Note: state ε-closure consist of final state of ε-NFA, the particular state is a final state in NFA.

$\hat{\delta}(q_0, \varepsilon) = \varepsilon\text{-closure}(q_0) = \{q_0q_1\}$
$\hat{\delta}(q_1, \varepsilon) = \varepsilon\text{-closure}(q_1) = q_1$
$\hat{\delta}(q_2, \varepsilon) = \varepsilon\text{-closure}(q_2) = \{q_0q_1q_2\}$

$\hat{\delta}(q_0, 0) = \varepsilon\text{-closure}(\delta(\hat{\delta}(q_0, \varepsilon), 0)$.

$= \varepsilon\text{-closure}(\delta(\{q_0q_1\}, 0)$.

$= \varepsilon\text{-closure}(\delta(q_0, 0) \cup \delta(q_1, 0))$.

$= \varepsilon\text{-closure}(q_0 \cup q_2)$.

$= \varepsilon\text{-closure}(q_0) \cup \varepsilon\text{-closure}(q_2)$.

$= \{q_0q_1\} \cup \{q_0q_1q_2\}$

$= \{q_0q_1q_2\}$

$\hat{\delta}(q_0, 1) = \varepsilon\text{-closure}(\delta(\hat{\delta}(q_0, \varepsilon), 1)$.

$= \varepsilon\text{-closure}(\delta(\{q_0q_1\}, 1)$.

$= \varepsilon\text{-closure}(\delta(q_0, 1) \cup \delta(q_1, 1))$.

=ε-closure(q_0 U ϕ).

=ε-closure(q_0) U ε-closure(ϕ).

={q_0q_1}

$\hat{\delta}$ (q_1, 0)= ε-closure(δ($\hat{\delta}$ (q_1, ε), 0).

= ε-closure(δ(q_1, 0)).

= ε-closure(q_2).

={$q_0q_1q_2$}

$\hat{\delta}$ (q_1, 1)= ε-closure(δ($\hat{\delta}$ (q_1, ε), 1).

= ε-closure(δ(q_1, 1)).

= ε-closure(ϕ).

=ϕ

$\hat{\delta}$ (q_2, 0)= ε-closure(δ($\hat{\delta}$ (q_2, ε), 0).

= ε-closure(δ({$q_0q_1q_2$}, 0).

= ε-closure(δ(q_0, 0) U δ(q_1, 0)U δ(q_2, 0)).

=ε-closure(q_0 U q_2 U ϕ).

=ε-closure(q_0) U ε-closure(q_2) U ε-closure(ϕ).

={q_0q_1} U {$q_0q_1q_2$} U (ϕ)

={$q_0q_1q_2$}

$\hat{\delta}$ (q_2, 1)= ε-closure(δ($\hat{\delta}$ (q_2, ε), 1).

= ε-closure(δ({$q_0q_1q_2$}, 1).

= ε-closure(δ(q_0, 1) U δ(q_1, 1) U δ(q_2, 1)).

=ε-closure(q_0 U ϕ U ϕ).

=ε-closure(q_0) U ε-closure(ϕ) U ε-closure(ϕ)

=(q_0 q_1) U (ϕ) U (ϕ)

= {q_0q_1}

Table 1.5: Equivalence table of ε-NFA and NFA.

Q \ Σ	0	1
→q_0	$\{q_0q_1q_2\}$	$\{q_0q_1\}$
q_1	$\{q_0q_1q_2\}$	φ
*q_2	$\{q_0q_1q_2\}$	$\{q_0q_1\}$

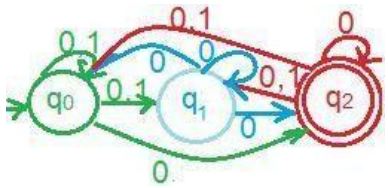


Figure 1.58: NFA without ε-transition.

Conversion of ε-NFA to DFA:

It is same as conversion of NFA to DFA, begins with initial state of ε-NFA.

Example 1.23: Convert the ε-NFA shown in figure1.59 into DFA.

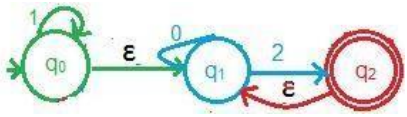


Figure 1.59: ε-NFA of Example 1.23.

First find out the ε-closure of all the states

$\hat{\delta}(q_0, \varepsilon)$ = ε-closure(q_0)= $\{q_0q_1\}$
$\hat{\delta}(q_1, \varepsilon)$ = ε-closure(q_1)= q_1
$\hat{\delta}(q_2, \varepsilon)$ = ε-closure(q_2) = $\{q_1q_2\}$

Table 1.6: Equivalence table of ε-NFA and NFA.

Q \ Σ	0	1	2
→q_0	q_1	$[q_0q_1]$	$[q_1q_2]$
q_1	q_1	φ	$[q_1q_2]$
$[q_0q_1]$	q_1	$[q_0q_1]$	$[q_1q_2]$
*$[q_1q_2]$	q_1	φ	$[q_1q_2]$

$\hat{\delta}(q_0, 0)$= ε-closure(δ($\hat{\delta}(q_0, \varepsilon)$, 0).

= ε-closure(δ($\{q_0q_1\}$, 0).

= ε-closure(δ(q_0, 0) U δ(q_1, 0)).

=ε-closure(ϕ U q_1).

=ε-closure(ϕ) U ε-closure(q_1).

= ϕ U q_1

=q_1

$\hat{\delta}$(q_0, 1)= ε-closure(δ($\hat{\delta}$(q_0, ε), 1).

= ε-closure(δ({q_0q_1}, 1).

= ε-closure(δ(q_0, 1)U δ(q_1, 1)).

=ε-closure(q_0U ϕ).

=ε-closure(q_0) U ε-closure(ϕ).

={q_0q_1}U ϕ

=[q_0q_1]

$\hat{\delta}$(q_0, 2)= ε-closure(δ($\hat{\delta}$(q_0, ε), 2).

= ε-closure(δ({q_0q_1}, 2).

= ε-closure(δ(q_0, 2)U δ(q_1, 2)).

=ε-closure(ϕ U q_2).

=ε-closure(q_2).

=[q_1q_2]

$\hat{\delta}$(q_1, 0)= ε-closure(δ($\hat{\delta}$(q_1, ε), 0).

= ε-closure(δ(q_1, 0).

= ε-closure(q_1).

= q_1

$\hat{\delta}$(q_1, 1)= ε-closure(δ($\hat{\delta}$(q_1, ε), 1).

= ε-closure(δ(q_1, 1).

= ε-closure(ϕ).

= ϕ

$\hat{\delta}(q_1, 2)$= ε-closure(δ($\hat{\delta}(q_1, \varepsilon)$, 2).

= ε-closure(δ(q_1, 2)

= ε-closure(q_2)

= [q_1 q_2]

$\hat{\delta}([q_0q_1], 0)$= ε-closure(δ($\hat{\delta}([q_0q_1], \varepsilon)$, 0)

= ε-closure(δ($\hat{\delta}(q_0, \varepsilon)$U $\hat{\delta}(q_1, \varepsilon)$), 0).

= ε-closure(δ((q_0, q_1) U q_1), 0)).

= ε-closure(δ((q_0, q_1), 0)).

= ε-closure(δ(q_0, 0) U δ(q_1, 0)).

= ε-closure(ϕ U q_1)

= ε-closure(q_1)

= q_1

$\hat{\delta}([q_0q_1], 1)$= ε-closure(δ($\hat{\delta}([q_0q_1], \varepsilon)$, 1)

= ε-closure(δ($\hat{\delta}(q_0, \varepsilon)$U $\hat{\delta}(q_1, \varepsilon)$), 1).

= ε-closure(δ((q_0, q_1) U q_1), 1)).

= ε-closure(δ((q_0, q_1), 1)).

= ε-closure(δ(q_0, 1) U δ(q_1, 1)).

= ε-closure(q_0 U ϕ)

= ε-closure(q_0)

= [q_0, q_1]

$\hat{\delta}([q_0q_1], 2)$= ε-closure(δ($\hat{\delta}([q_0q_1], \varepsilon)$, 2)

= ε-closure(δ($\hat{\delta}(q_0, \varepsilon)$U $\hat{\delta}(q_1, \varepsilon)$), 2).

= ε-closure(δ((q_0, q_1) U q_1), 2)).

= ε-closure(δ((q_0, q_1), 2)).

= ε-closure(δ(q_0, 2) U δ(q_1, 2)).

= ε-closure(ϕ U q_2)

= ε-closure(q_2)

= [q_1 q_2]

$\hat{\delta}$([q_1q_2], 0)= ε-closure(δ($\hat{\delta}$([q_1q_2], ε), 0)

= ε-closure(δ($\hat{\delta}$(q_1, ε)U $\hat{\delta}$(q_2, ε), 0)).

= ε-closure(δ(q_1U (q_1 q_2)), 0).

= ε-closure(δ(q_1, q_2), 0)).

= ε-closure(δ(q_1, 0) U δ (q_2, 0))

= ε-closure(q_1 U ϕ).

= ε-closure(q_1)

= q_1

$\hat{\delta}$([q_1q_2], 1)= ε-closure(δ($\hat{\delta}$([q_1q_2], ε), 1)

= ε-closure(δ($\hat{\delta}$(q_1, ε)U $\hat{\delta}$(q_2, ε), 1)).

= ε-closure(δ(q_1U (q_1 q_2)), 1).

= ε-closure(δ(q_1, q_2), 1)).

= ε-closure(δ(q_1, 1)U δ (q_2, 1))

= ε-closure(ϕ U ϕ).

= ε-closure(ϕ)

= ϕ

$\hat{\delta}$([q_1q_2], 2)= ε-closure(δ($\hat{\delta}$([q_1q_2], ε), 2)

= ε-closure(δ($\hat{\delta}$(q_1, ε)U $\hat{\delta}$(q_2, ε), 2)).

= ε-closure(δ(q_1U (q_1 q_2)), 2).

= ε-closure(δ(q_1, q_2), 2)).

= ε-closure(δ(q_1, 2)U δ (q_2, 2))

= ε-closure(q_2 U ϕ).

= ε-closure(q_2)

= [q_1 q_2]

Table 1.7: Equivalence table of ε-NFA and DFA.

Q / Σ	0	1	2
q_0	q_1	[q_0q_1]	[q_1q_2]
q_1	q_1	φ	[q_1q_2]
[q_0 q_1]	q_1	[q_0 q_1]	[q_1 q_2]
[q_1q_2]	q_1	φ	[q_1 q_2]
φ	φ	φ	φ

$\hat{\delta}$ (ϕ, (0,1,2))= ϕ

Mark the final states, any pair which has a final state of ε -NFA the particular pair is afinal state in DFA.

Table 1.7: Equivalence table of ε-NFA and DFA .

Q / Σ	0	1	2
q_0	q_1	[q_0q_1]	[q_1q_2]
q_1	q_1	φ	[q_1q_2]
[q_0 q_1]	q_1	[q_0 q_1]	[q_1 q_2]
*[q_1q_2]	q_1	φ	[q_1 q_2]
φ	φ	φ	φ

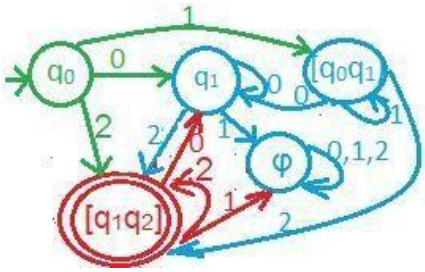


Figure 1.60: DFA of figure 1.59.

MODULE 2

REGULAR EXPRESSIONS

Language represented in the form of expression (symbols separated by operator) is called regular expression.

$\sum$= {a, b, 0, 1}

Language L= $\sum$*= {a, b, 0, 1}* ={ Ɛ, a, b, 0, 1, aa, bb, 00, 11…}

Regular expression R= { Ɛ + a + 0 + 1 + aa + bb + 00 + 11…}

Before studying regular expression in detail, we will discuss some fundamental rules of the regular expression.

i. Φ+1=1 (φ is like, number 0).
ii. Ɛ + 1= Ɛ+1 (Ɛ is like number 1)
iii. φƐ=φ
iv. Ɛ+φ=Ɛ
v. 1+1=1
vi. 0+0=0
vii. 0+1=0+1
viii. 0+01=0(Ɛ+1)
ix. 0Ɛ=0
x. 0φ=0
xi. Φ*={Ɛ+ φ+ φφ…}={Ɛ+φ}=Ɛ
xii. 0^++Ɛ=0*
xiii. 0*+Ɛ=0*
xiv. $(1+Ɛ)^+$=1*
xv. $(1+φ)^+=1^+$
xvi. $00*=0^+$
xvii. $Ɛ+11^*=Ɛ+1^+=1*$
xviii. $(Ɛ+11)(Ɛ+11)*=(Ɛ+11)^+=(11)*$
xix. (00)*(1+01)=(00)*(Ɛ+0)1= (Ɛ+00+0000..)(Ɛ+0)1=0*1
xx. Write a regular expression set of strings made up of 0's and 1's ending with 00
RE=(0+1)*00.
xxi. Write a regular expression for all the set of string made up of 0's and 1's which begin with 0and end with 1
RE=0(1+0)*1
xxii. Strings begins with 0 or 1 is represented as $(0+1)(0+1)^*$
xxiii. All strings begins with 0 or 1 and not having two consecutive 0's
RE=(0+1)(1+10)
xxiv. Write regular expression which consist of two consecutive 0's
RE=(0+1)*00(0+1)*
xxv. All the strings of 0's and 1's whose last two symbols are same
RE=(0+1)*(00+11)
xxvi. Sets of all strings which starts with either 00 or 11 or ends with 00 or 11

RE=(00+11) (0+1)* +(0+1)*(00+11).

xxvii. Find the regular expression of the language L={$a^n b^m$:n≥0, m≥0}
L={Ɛ+a+aa..}{Ɛ+b+bb…}
RE=a*b*

xxviii. Find the regular expression of the language L={$a^n b^{2m+2}$: n≥0, m≥0}
RE=(aa)*(bb)*bb.

xxix. Find the regular expression of the language L={$a^{2n} b^m$: n≥4, m≤3}
RE=(aaaa)*(Ɛ+b+bb+bbb)

xxx. Find the regular expression of the language L={$a^n b^m$/n+m is even}
RE=(aa)*(bb)*+a(aa)*b(bb)*

Conversion from Regular expression to finite automata:

Three basic regular expressions are used in conversion (0+1), 01 and 0*.

The finite automata which accept either 0 or 1 is shown in figure 2.1.

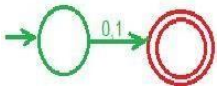


Figure 2.1: Finite automata

But in regular expression each symbol is having its own finite automata.

Definition for the finite automat which accept the symbol '1' is

$M_1 = (Q_1, \sum_1, \delta, IS_1, FS_1)$

Q_1={C, D}; $\sum_1$=1; δ(C, 1) = D; IS=C; FS=D

Definition for the finite automat which accept the symbol '0' or '1' is

M= (Q,∑, δ, IS, FS)

Q={Q0 U Q1 U E U F} ={ A, B, C, D, E. F);

∑= {0, 1}

δ

δ(A, 0)= B; δ (C, 1)= D; δ(E, Ɛ)={A, C}; δ(D, Ɛ)=F: δ(D, Ɛ)=F

IS=E; FS=F

For the symbol '0' the finite automata is shown in figure 2.2

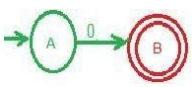


Figure 2.2: Finite automata for regular expression 0.

Definition for the finite automat which accept the symbol ‘0’ is

$M_0= (Q_0,\sum_0, \delta, IS_0, FS_0)$

$Q_0=\{A, B\}; \sum_0=0; \delta(A, 0)= B; IS=A; FS=B$

For the symbol ‘1’ is shown in figure 2.3

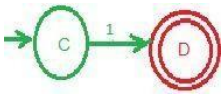


Figure 2.3: Finite automata for regular expression 1.

Definition for the finite automat which accept the symbol ‘1’ is

$M_1= (Q_1, \sum_1, \delta, IS_1, FS_1)$

$Q_1=\{C, D\}; \sum_1=1; \delta(C, 1)= D; IS=C; FS=D$

We are joining these two finite automata with disturbing the regular expression 0+1 as shown in figure 2.4. To join these two finite automata we need more extra states, one for the initial state and second is for the final state, states A and B no more initial states and B and D are no more final states. All these states will now become intermediate states as shown in figure 2.4.

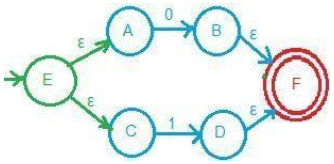


Figure 2.4: Finite automata for regular expression 0+1.

Definition for the finite automat which accept the symbol ‘0’ or ‘1’ is

$M= (Q,\sum, \delta, IS, FS)$

$Q=\{Q_0 \cup Q_1 \cup E \cup F\} =\{ A, B, C, D, E. F);$

$\sum=\{0, 1\}$

δ

$\delta(A, 0)= B; \delta (C, 1)= D; \delta(E, Ɛ)=\{A, C\}; \delta(D, Ɛ)=F: \delta(D, Ɛ)=F$

IS=E; FS=F

For the regular expression 01 the finite automata is shown in figure 2.6.

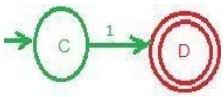


Figure 2.5: Finite automata for regular expression 0 and 1.

Rewrite the expression 01 as 0Ɛ1.

We can join the two finite automata shown in figure2.5 and figure 2.6 without disturbing the regular expression 01 is shown in figure 2.6.

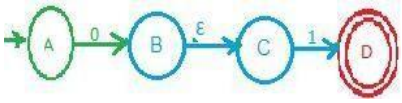


Figure 2.6: Finite automata for regular expression 01.

Definition for the finite automat which accept the string 01 is

M= (Q,∑, δ, IS, FS)

Q={Q_0 U Q_1} ={ A, B, C, D);

∑={0, 1}

δ

δ(A, 0)= B; δ (C, 1)= D; δ(B, Ɛ)= C

IS=A; FS=D

Like that we can design the finite automata for 0*.

0*= {Ɛ, 0, 00, 000…}

The regular expression for symbol '0' is shown in figure 2.7

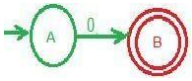


Figure 2.7: Finite automata for regular expression 0.

The designed finite automata accept all the strings of 0* ie., {Ɛ, 0, 00, 000…}

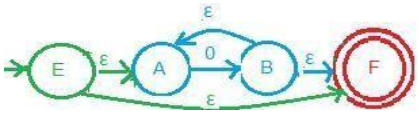


Figure 2.8: Finite automata for regular expression 0*.

Definition for the finite automat which accept the string 0* is

M= (Q,∑, δ, IS, FS)

Q={Q_0 U E U F} ={ A, B, E, F);

∑=0

δ

δ(A, 0)= B; δ (E, Ɛ)= {A, F}; δ(B, Ɛ)= {F, A}

IS=E; FS=F

Example 2.1: Design a finite automat for the regular expression 0*1

i. Draw the finite automata for the symbol '0'

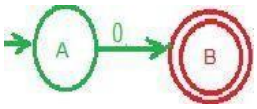


Figure 2.9: Finite automata for regular expression 0.

ii. Convert the finite automata of symbol '0' to 0*

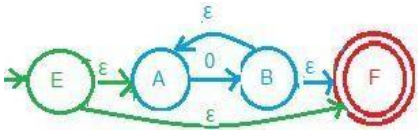


Figure 2.10: Finite automata for regular expression 0*.

iii. Draw the finite automata for the symbol '1'

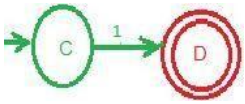


Figure 2.11: Finite automata for regular expression 1.

Join finite automata of 0* and 1, which will be the solution for the regular expression 0*1

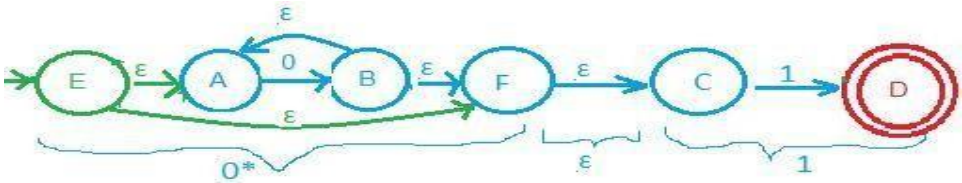


Figure 2.12: Finite automata for regular expression 0*1.

Example 2.2: Design a finite automat for the regular expression 0(0+1)*1

i. Draw the finite automata for the symbol '0' and '1'

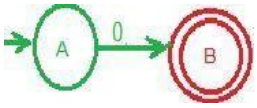


Figure 2.13: Finite automata for regular expression 0.

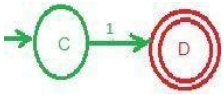


Figure 2.13: Finite automata for regular expression 1.

ii. Construct the finite automata for the expression 0 + 1

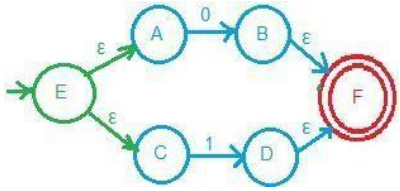


Figure 2.14: Finite automata for regular expression 0+1.

iii. Construct the finite automata for the symbol (0 + 1)*

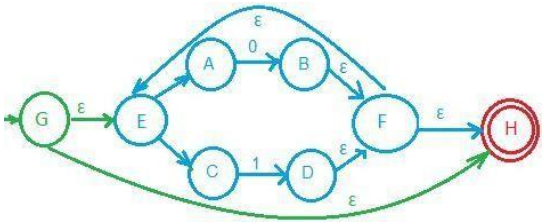


Figure 2.15: Finite automata for regular expression (0+1)*.

iv. Join the finite automata of '0', (0+1)* and '1'

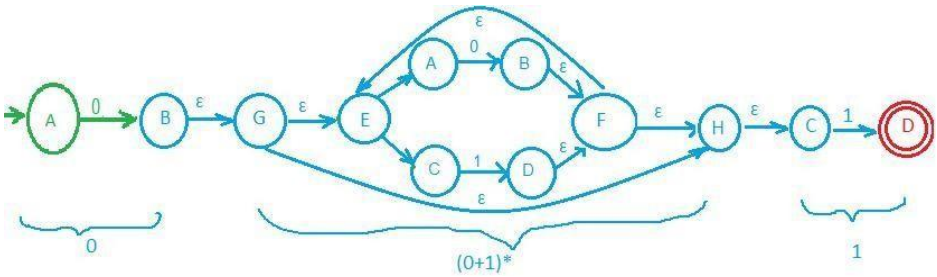


Figure 2.16: Finite automata for regular expression 0(0+1)*1.

Conversion from Finite Automata to Regular Expression:

They are of two methods of conversions, they are:

i. Recursive Method.
ii. State Elimination Method.

i. Recursive Method :

In this method it uses recursive method, to construct new regular expression.
For example the regular expression of DFA shown in figure 2.17 is

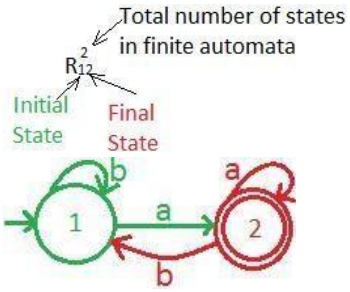


Figure 2.17: A typical Finite automata.

In order to obtain the value of regular expression R_{12}^2 , we use all the (transition) possible regular expression of the DFA. The possible regular expression of the figure 2.17 is shown in the table2.1.

Table 2.1: Recursive table

	k=0	k=1
R_{11}^k		
R_{12}^k		
R_{21}^k		
R_{22}^k		

Equations to obtain the regular expression of all the transitions are as follows:

if k=0 $R_{ij}^k = \begin{cases} a \text{ if } i \neq j \\ a \cup \varepsilon \text{ if } i = j \end{cases}$

where a is the transition between state i and j

if k>0 $R_{ij}^k = R_{ik}^{k-1}(R_{kk}^{k-1})^* R_{kj}^{k-1} + R_{ij}^{k-1}$

If state 2 and 3 are final states then the regular expression is written as:

$RE = R_{12}^3 + R_{13}^3$

Example 2.3: Find the regular expression of DFA shown in figure 2.17 by using recursive method.

	k=0	k=1
R_{11}^k	$b+\varepsilon$	b^*
R_{12}^k	a	b^*a
R_{21}^k	b	b^*b
R_{22}^k	$a+\varepsilon$	$b^*a+\varepsilon$

$R_{11}^1 = R_{11}^0(R_{11}^0)^* R_{11}^0 + R_{11}^0$

$R_{11}^1 = (b+\varepsilon)(b+\varepsilon)^*(b+\varepsilon) + (b+\varepsilon)$

$= [(b+\varepsilon)^+ + \varepsilon](b+\varepsilon)$

$= [(b+\varepsilon)^*(b+\varepsilon)$

$= [(b+\varepsilon)^+$

$= b^*$

$R_{12}^1 = R_{11}^0(R_{11}^0)^* R_{12}^0 + R_{12}^0$

$= (b+\varepsilon)(b+\varepsilon)^* a + a$

$= [(b+\varepsilon)(b+\varepsilon)^* + \varepsilon]a$

$= [(b+\varepsilon)^+ + \varepsilon]a$

$= b^*a$

$R_{21}^1 = R_{21}^0(R_{11}^0)^* R_{11}^0 + R_{21}^0$

$= b(b+\varepsilon)^*(b+\varepsilon) + b$

$= [(b+\varepsilon)^*(b+\varepsilon) + \varepsilon]b$

$= [(b+\varepsilon)^+ + \varepsilon]b$

$= (b+\varepsilon)^+ b$

$= b^*b$

$R_{22}^1 = R_{21}^0(R_{11}^0)^* R_{12}^0 + R_{22}^0$

$= b\,b^*a + a + \varepsilon = b^+a + a + \varepsilon$

$= (b^+ + \varepsilon)a + \varepsilon$

$= b^*a + \varepsilon$

$R_{12}^2 = R_{12}^1(R_{22}^1)^* R_{22}^1 + R_{12}^1$

$RE = b^*a(b^*a+\varepsilon)^*(b^*a+\varepsilon) + b^*a$

Example 2.4: Find the regular expression of DFA shown in figure 2.18 by using recursive method.

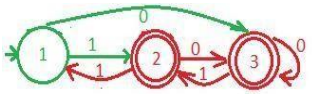


Figure 2.18: DFA

	k=0	k-1	k=2
R_{11}^k	φ+ε =ε	ε	(11)*
R_{12}^k	1	1	1(11)*
R_{13}^k	0	0	1*0
R_{21}^k	1	1	(11)*1
R_{22}^k	ε	11+ε	(11)*
R_{23}^k	0	10+0	(11)*0(1+ε)
R_{31}^k	φ	φ	11(11)*
R_{32}^k	1	1	1(11)*
R_{33}^k	0+ε	0+ε	101*+0+ε

$R_{11}^1=R_{11}^0(R_{11}^0)^*R_{11}^0+R_{11}^0$

$=\varepsilon(\varepsilon)^*\varepsilon+\varepsilon=\varepsilon$

$R_{12}^1=R_{11}^0(R_{11}^0)^*R_{12}^0+R_{12}^0$

$=\varepsilon(\varepsilon^*)1+1=1$

$R_{12}^1=R_{11}^0(R_{11}^0)^*R_{12}^0+R_{12}^0$

$=\varepsilon(\varepsilon^*)1+1=1$

$R_{13}^1=R_{11}^0(R_{11}^0)^*R_{13}^0+R_{13}^0$

$=\varepsilon(\varepsilon)^*0+0$

$=0+0=0$

$R_{21}^1=R_{21}^0(R_{11}^0)^*R_{11}^0+R_{21}^0$

$=1(\varepsilon)\varepsilon+1=1$

$R_{22}^1=R_{21}^0(R_{11}^0)^*R_{12}^0+R_{22}^0$

$=1(\varepsilon^*)1+\varepsilon$

$=11+\varepsilon$

$R_{23}^1=R_{21}^0(R_{11}^0)^*R_{13}^0+R_{23}^0$

$=1(\varepsilon)^*0+0$

$=10+0$

$R_{31}^1=R_{31}^0(R_{11}^0)^*R_{11}^0+R_{31}^0$

$=\varphi+\varphi=\varphi$

$R_{32}^1=R_{31}^0(R_{11}^0)^*R_{12}^0+R_{32}^0$

$=\varphi+1=1$

$R_{33}^1=R_{31}^0(R_{11}^0)^*R_{13}^0+R_{33}^0$

$=\varphi+0+\varepsilon$

$R_{11}^2=R_{12}^1(R_{22}^1)^*R_{21}^1+R_{11}^1$

$1(11+\varepsilon)^*1+\varepsilon$

$=11(11)^*+\varepsilon$

$=(11)^+ +\varepsilon$

$=(11)^*$

$R_{12}^2=R_{12}^1(R_{22}^1)^*R_{22}^1+R_{12}^1$

$=1(11+\varepsilon)^*(11+\varepsilon)+1$

$=1(11+\varepsilon)^+1$

$=1[(11+\varepsilon)^+ +\varepsilon]$

$=1(11)^*$

$R_{13}^2=R_{12}^1(R_{22}^1)^*R_{23}^1+R_{13}^1$

$=1(11+\varepsilon)^*(10+0)+0$

$=1(11)^*(1+\varepsilon)0+0$

$=[1(11)^*+1(11)^*]0+0$

$=[(11)^+ +1(11)^*+\varepsilon]0$

$=[(11)^*+1(11)^*]0$

$=(11)^*(\varepsilon+1)0$ $=1^*$

$=1^*0$

$R_{21}^2=R_{22}^1(R_{22}^1)^*R_{21}^1+R_{21}^1$

$=(11+\varepsilon)(11+\varepsilon)^*1+1$

$=(11+\varepsilon)^+1+1$

$=((11+\varepsilon)^+ +\varepsilon)1$

$=(11)^*1$

$R_{22}^2=R_{12}^1(R_{22}^1)^*R_{22}^1+R_{22}^1$

$=(11+\varepsilon)(11+\varepsilon)^*(11+\varepsilon)+(11+\varepsilon)$

$=((11+\varepsilon)(11+\varepsilon)^*+\varepsilon)\ (11+\varepsilon)$

$=((11+\varepsilon)^+ +\varepsilon)\ (11+\varepsilon)$

$=((11+\varepsilon)^*(11+\varepsilon))$

$=((11+\varepsilon))^+$

$=(11)^*$

$R_{23}^2=R_{22}^1(R_{22}^1)^*R_{23}^1+R_{23}^1$

$=(11+\varepsilon)((11+\varepsilon)^*(10+0)+(10+0)$

$=[(11+\varepsilon)((11+\varepsilon)^*+\varepsilon)](10+0)$

$=[(11+\varepsilon)^+ +\varepsilon]\ (10+0)$

$=(11+\varepsilon)^*(10+0)$

$=(11)^*0(1+\varepsilon)$

$R_{31}^2=R_{32}^1(R_{22}^1)^*R_{21}^1+R_{31}^1$

$=1(11+\varepsilon)^*1+\varphi$

$=11(11)^*$

$R_{32}^2=R_{32}^1(R_{22}^1)^*R_{22}^1+R_{32}^1$

$=1(11+\varepsilon)^*(11+\varepsilon)+1$

$=1((11+\varepsilon)^+1$

$=1[(11+\varepsilon)^+ +(\varepsilon)]$

$=1(11+\varepsilon)^*$

$=1(11)^*$

$R_{33}^2=R_{32}^1(R_{22}^1)R_{23}^1+R_{33}^1$

$=1(11+\varepsilon)^*(10+0)+(0+\varepsilon)$

$=10(11+\varepsilon)^*(1+\varepsilon)+(0+\varepsilon)$

$=10(11)^*(1+\varepsilon)+(0+\varepsilon)$

$=101^*+0+\varepsilon$

$RE=R_{12}^3+R_{13}^3$

$=[1^*00(11^*)^*]1(11)^*+1(11)^*\ (101^*+0+\varepsilon)^*(101^*+0+\varepsilon)+1^*0$

State Elimination Method:

In state elimination method, we obtain regular expression, by eliminating all the intermediate states. After eliminating the intermediate states, it is easy to find the regular expression for the remaining states (initial and final).

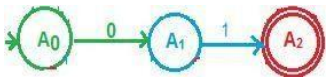


Figure2.19: Finite automata

The regular expression of the finite automata shown in figure2.20 is:

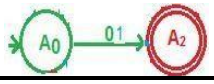


Figure2.20: Finite automata

RE=01

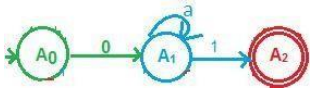


Figure2.21: Finite automata

The regular expression of the finite automata shown in figure 2.2 is:

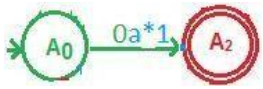


Figure2.22: Finite automata

RE=0a*1

From the above examples we come to know that eliminating state has one predecessor state A_0 and one successor state A_2. In complex finite automata the eliminating state may have more number of predecessors and successors. One Such example is shown in the figure 2.23.

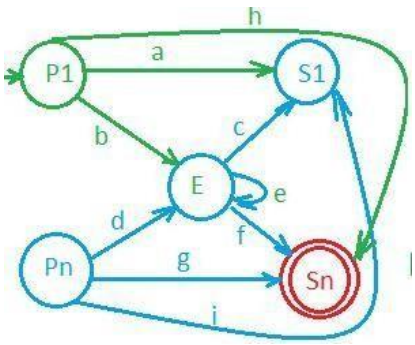


Figure 2.23: Eliminating state E has n predecessor and n successor.

Before eliminating any state from the finite automata, we need to find out the path passing through the eliminating state. The paths passing through the eliminating state are P1ESn, P1ES1, PnES1 and PnESn. After eliminating state E, the automata is shown in the figure 2.24.

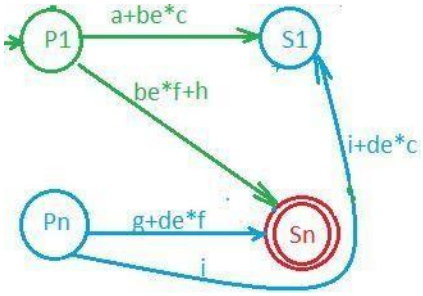


Figure 2.24: Finite automata after eliminating state E.

After eliminating all predecessors and successors, leaving initial state and final state, the remaining machine looks shown in the figure 2.25.

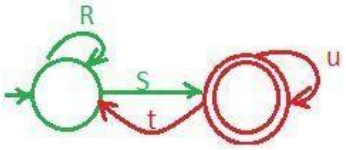


Figure 2.25: Finite automata after eliminating all predecessors and successors.

The regular expression is, language accepted by the initial state of finite automata and Su*.

The language accepted by the initial state is {R*+Su*t+R*Su*t + su*tR*…}, this is like {x*+y*+xy + yx …}

We can simplify it as (x+y)* similarly, we can simplify the language for the initial state as (R+Su*t)*. Therefore, regular expression of finite automata is (R+Su*t)*Su*. Note that in some cases the same state is initial as well as final state.

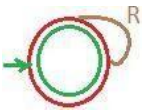


Figure 2.26: Same state is initial as well as final state.

The regular expression is R*.

To eliminate state, three rules are used, they are:

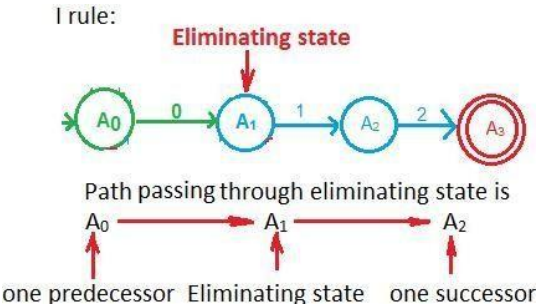


Figure 2.27: One path passing through eliminating state A_1

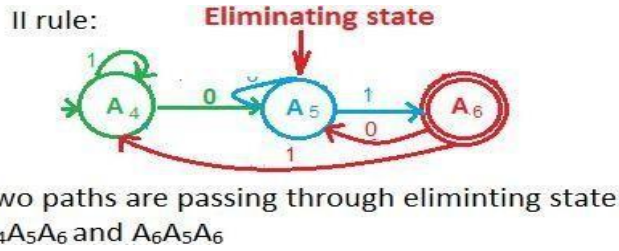


Figure 2.28: Two paths are passing through eliminating state A_1

Predecessor state	Eliminating State	Successor state
A_4	A_5	A_6
A_6	A_5	A_6

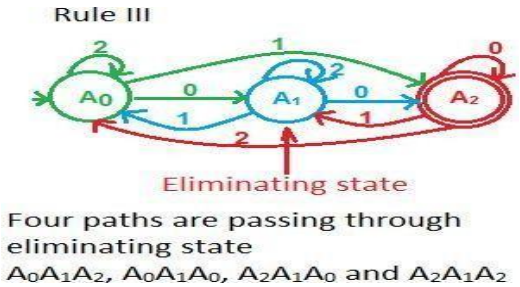


Figure 2.29: Four paths are passing through eliminating state A_1

Predecessor state	Eliminating State	Successor state
A_0	A_1	A_2
A_0	A_1	A_0
A_2	A_1	A_0
A_2	A_1	A_2

Example 2.5: Construct regular expression for the finite automat shown in figure 2.30.

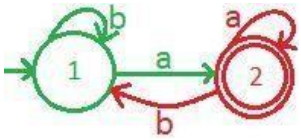


Figure 2.30: Finite Automata

No intermediate in the given finite automata, hence we can't eliminate any of the state.

R=b; S=a; u=a; t=b;

Therefore RE =(R+Su*t)*Su*

= (b+aa*b)*aa*

=b(Ɛ+aa*)*aa*

=b(aa*)*aa*

Example 2.6: Construct regular expression for the finite automat shown in figure 2.31 by using state elimination method.

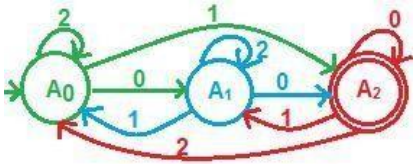


Figure 2.31: Finite Automata.

We can eliminate one intermediate state A_1, the paths passing through the eliminating state are $A_0A_1A_0$, $A_0A_1A_2$, $A_2A_1A_0$ and $A_2A_1A_2$.

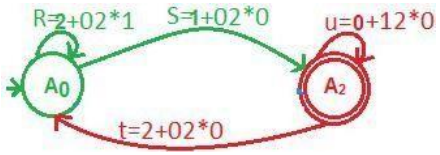


Figure 2.32: Final finite Automata

RE=(R+Su*t)*Su*

= ((2+02*1)+(1+02*0)(0+12*0)*(2+02*0))*((1+02*0)(0+12*0)*.

Application of Regular expression:

Pumping lemma:

It is used to prove language as non-regular for example

L=(00, 000, 0000…}

The automat is shown in figure 2.33 which accept 00.

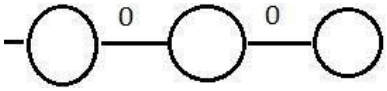


Figure 2.33: Finite Automata

The automat is shown in figure 2.34 which accept 000.

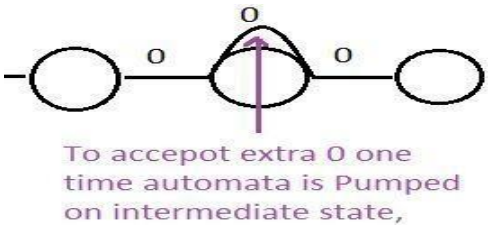


Figure 2.34: Finite Automata pumped at intermediate state one time.

The automat is shown in figure 2.35 accept 0000.

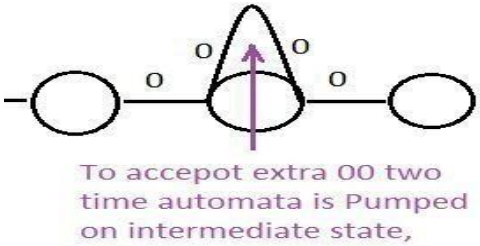


Figure 2.35: Finite Automata pumped at intermediate state one time.

If more 0s, there will be more number of pumping. In such a case it is very difficult to count the number of pumping. To count the number of pump easily we draw circles around the intermediate state so that we can easily count the number of pumping to prove the language as non-regular.

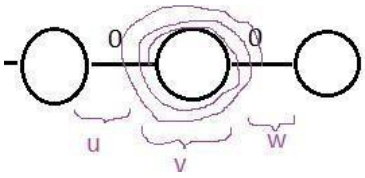


Figure 2.35: Finite Automata pumped at intermediate state finite time.

We can write $uv^iw \in L$ for i={1, 2, 3...}.

To prove the language as non-regular, smallest string of language is divided into uvw by considering two conditions.

i. $V \notin \varepsilon$.
ii. $|UV| \leq n$ where n is the smallest string of the language. To prove the language as a non-regular, throughout the proof the value of n is constant.

Example 2.7: Prove that the language $L=\{0^{i^2} : i \geq 1\}$ is not regular.

L={0, 0000, 000000000...}
n u v w

$uv^iw \in L$ for i={1, 2…}

$0(00)^i000000 \in L$

I=1

000000000 ЄL

I=2

00000000000 $\notin$ L

When i=2 the string obtained is not belongs to language, hence proved language is not regular.

Example 2.8: Prove that the language $L=\{a^nb^{n+1} : n \geq 2\}$ is not regular.

L={aabbb, aaabbbb....}
n u v w

i=1

aaabbbb Є L

I=2

Aaabaabbbb $\notin$ L

When i=2 the string obtained is not belongs to language, hence proved the language is non-regular.

Closure Properties of Regular Expression:

Till now we have seen many regular expressions and their applications. These regular expressions are bounded/ related by operators called closure properties of regular expression. The different operators used to relate regular expressions are:

i. Boolean.
ii. Reversal.
iii. String Homomorphism and Inverse Homomorphism.

i. Boolean Operator:

They are of different types

a. Union b. Intersection c. Difference d. Compliment.

a. Union Operation.

To perform union operation it uses following steps.

- This operation is applied on two machines defined as

$M_1=(Q_1, \sum_1, \delta_1, IS_1, FS_1)$

$M_2=(Q_2, \sum_2, \delta_2, IS_2, FS_2)$

- Find the language L1 and L2 of the two machines.
- Perform union operation on the two languages L1 and L2 to obtain language L3
- Combine two machines to perform union operation.
- Apply language L3 on combined machines to obtain machine for union operation.

Same steps are used for the other Boolean operation.

$M_1UM_2=(Q_1UQ_2, \sum_1U\sum_2, \delta_1U\delta_2, ?, ?)$

For example consider the machine 1 is shown in figure 2.36

i.

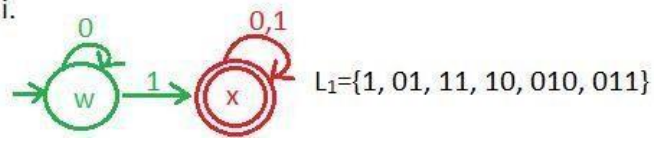


Figure 2.36: Finite Automata

Machine 2 is shown in figure 2.37

ii.

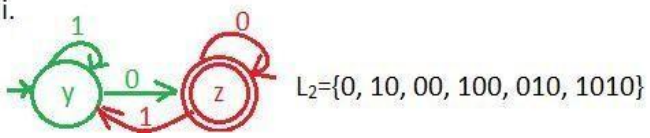


Figure 2.37: Finite Automata

The union operation on two languages is

$L_1UL_2=\{0, 1, 00, 01, 10, 11, 010, 011, 010, 100, 1010\}$

In order to obtain the finite automata for the union operation, we need to join two finite automata. The transition table for union operation is shown in table…

Table 2.1: Transition table for Boolean operation

δ_{1U2}	0	1
w, y	w, z	x, y
w, z	w, z	x, y
x, y	x, z	x, y
x, z	x, z	x, y

Equivalent diagram is shown in the figure 2.38

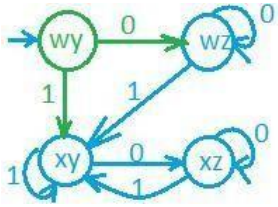


Figure 2.38: Finite Automata after combining figure 2.36 and 2.37.

According to union operation it has to accept 0, 00

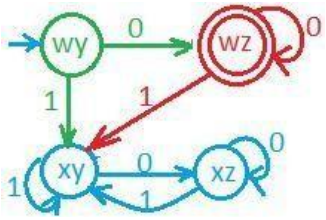


Figure 2.39: Finite Automata to accept 0, 00.

it has to accepts 1, 01, 11, 011

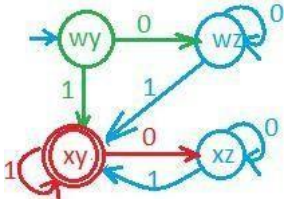


Figure 2.40: Finite Automata accepts 1, 01, 11, 011.

It also accepts 10, 010…

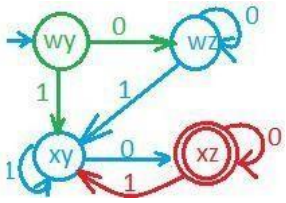


Figure 2.41: Finite Automata accepts 10, 010...

The final automata for the union operation is shown in figure 2.42

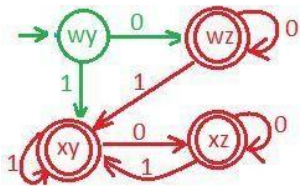


Figure 2.42: Union of two finite automata of figure 2.36 and 2.37.

For intersection M1 ∩ 2={10, 100, 010, 1010…}

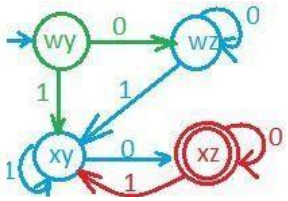


Figure 2.43: Intersection of two finite automata of figure 2.36 and 2.37.

For differentiation M_{1-2} ={1, 11, 01, 011, 101}

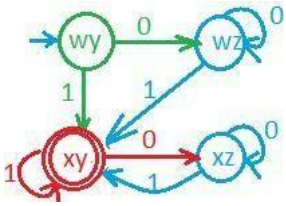


Figure 2.44: Difference of two finite automata of figure 2.36 and 2.37.

Compliment:

To compliment the regular expression it uses the following steps:

i. Obtain the Ɛ-NFA/NFA for the regular expression.
ii. Convert the Ɛ-NFA to DFA.
iii. Compliment DFA.
iv. Convert DFA to regular expression.

Example 2.9: Find the compliment of regular expression 0.

i. Obtain the Ɛ-NFA/NFA for the regular expression.

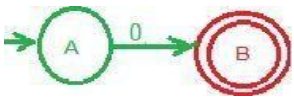


Figure 2.45: finite automata for regular expression 0.

ii. Convert the Ɛ-NFA/NFA to DFA.

δ(A, 0)=B; δ(B, 0)=φ; δ(φ,0)=φ

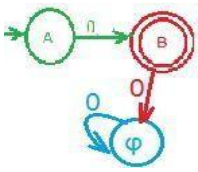


Figure 2.46 : Equivalent DFA for the figure 2.45.

iii. Compliment DFA. Reverse all transitions and convert all final to non final and vice-versa.

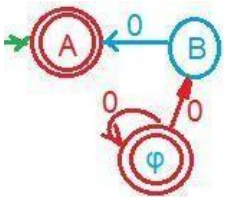


Figure 2.47: Complimented DFA of figure 2.46.

State B and φ are not reachable from the initial State. Hence the two states are removed from the finite automata.

Figure 2.48. Complimented DFA for the regular expression 0.

The regular expression of finite automata shown in figure 2.48 is

RE=Ɛ

Reversal of regular expression:

i. Reverse all transitions.
ii. Convert initial state to final and final state to intermediate state.
iii. Draw epsilon-transition between new initial state to all previous final states, mark new state as an initial state.

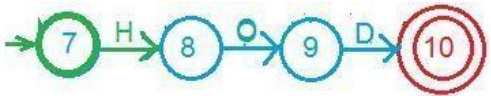


Figure 2.49: Non-deterministic finite automata.

Figure 2.50: Reversed NFA of the figure 2.49

String Homomorphism:

String homomorphism is a function on the language, it replaces string by symbol

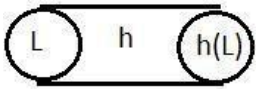


Figure 2.51: Homomorphism.

Inverse Homomorphism:

It is an inverse string homomorphism

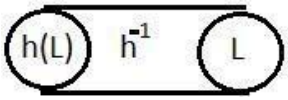


Figure 2.52: Inverse string homomorphism

Example 2.10: Consider the homomorphism, h(0)=abb; h(1)= a; h(011)=abbaa

If the homomorphism from an alphabet {0, 1, 2} to {a, b} is defined by h(0)=ab; h(1)=b; & h(2)=aa, then:

i. What is h(2201).
ii. If L is language 1*02* what is h(L).
iii. If L is language(ab+baa)*bab then what is $h^{-1}(L)$

i. h(2201)=aaaaabb
ii. h(L)=h(1*02*)=b*ab(aa)*
iii. (ab+baa)*bab=(0+12)*10

Equivalence and minimization:

In Equivalence, two different DFA's are given, we need to find out the two DFA's are similar or not. By filling table and Preshapat tree we can decide the two DFAs are similar | not similar.

Preshapat Tree: To find pair of nodes similar | not similar Preshapat tree is used. It consists of root node, intermediate nodes and leaves nodes. The root node of the tree is pair of initial states of two DFAs. Next we construct branches of tree depending on number of input symbols. If leaves node do not have a cross in the difference table and same pairs are repeating, we stop constructing tree and decide the root pair is similar or else dissimilar, any of leaves node has a cross in the difference table.

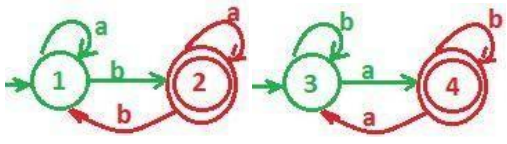


Figure 2.53: Deterministic finite automata.

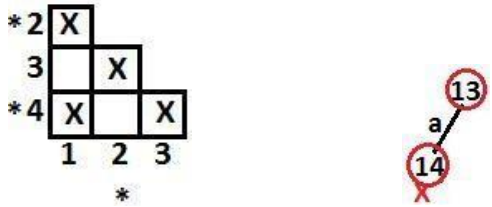


Figure 2.54: Difference table with Preshapat tree.

Leaves node 14 has cross in the difference table, so two DFA's are not equal.

For example

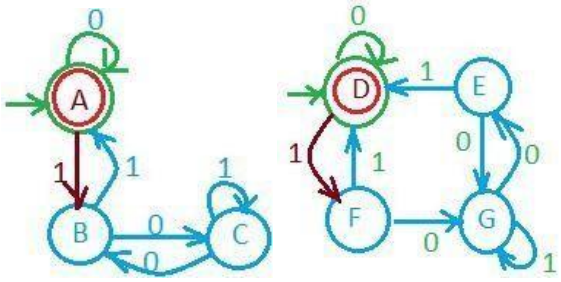


Figure 2.55: Deterministic finite automata.

Note that root of the tree is pair of initial states. Leaves nodes of the tree do not have cross in the difference table and nodes are repeating, hence the two DFA's are similar.

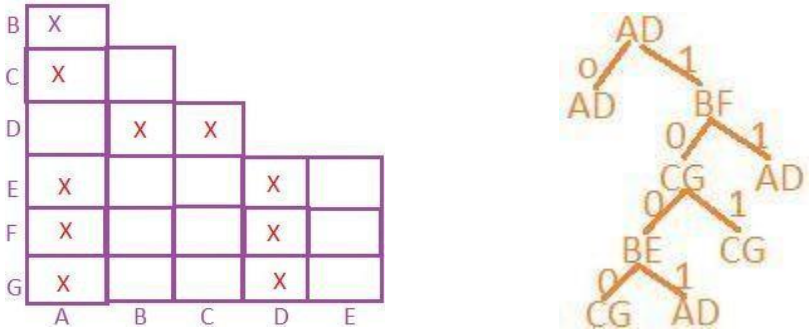


Figure 2.56: Difference table with Preshapat tree.

Example 2.11: Find out the two DFAs are similar or not.

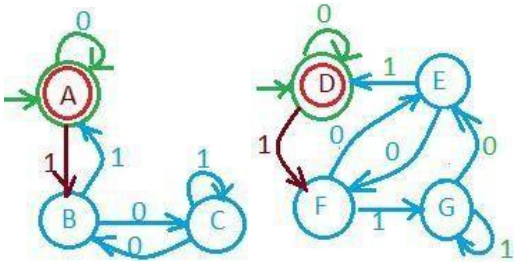


Figure 2.57: Deterministic finite automata.

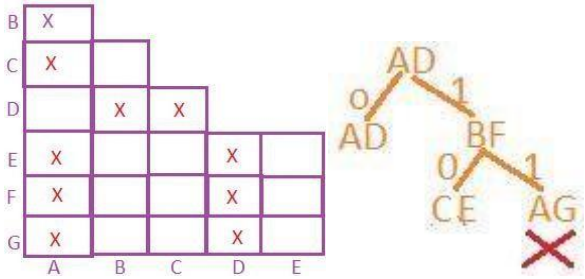


Figure 2.58: Transition tree with table.

The leaves node AG has X in the table; hence two DFAs are dissimilar.

In case of minimization, the two states are minimized to one state if they are similar. For example state X and Y are equal, they are replaced by one state as shown in figure 2.59

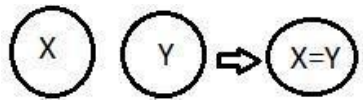


Figure 2.59: Similar states are replced into single state.

To decide the states are similar | not similar we use Shanprepat tree.

Shanprepat Tree: To find equivalence between pair of states Shanprepat tree is used. It consists of root node, intermediate nodes and leaves nodes. The root node of the tree is pair of states used to find the similar | not similar states. Next we construct branches of tree depending on number of input symbols. If leaves node do not have a cross in the difference table and same pairs are repeating, we stop constructing tree and decide the root pair is similar or else decide two states are dissimilar, any of leaves node has a cross in the difference table.

Example 2.12: Using table filling method, minimize the DFA shown in figure 2.60 and draw the transition diagram of the resulting DFA.

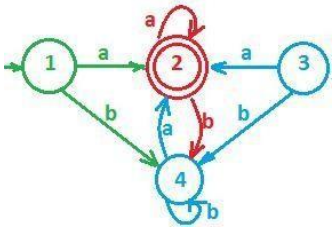


Figure 2.60: Deterministic finite automata.

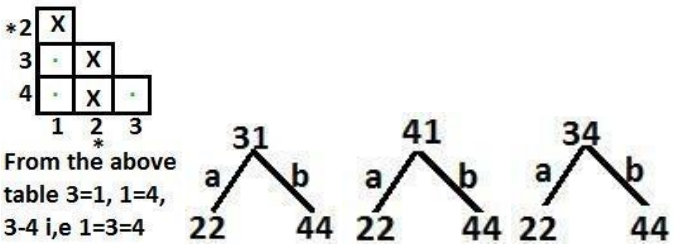


Figure 2.61: Differnce table with Shanprepat Tree

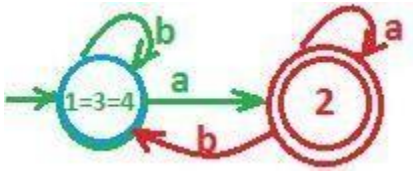


Figure 2.62: Minimized DFA of the figure 2.60

Example 2.13: Using table filling method, minimize the DFA shown in figure 2.63 and draw the transition diagram of the resulting DFA.

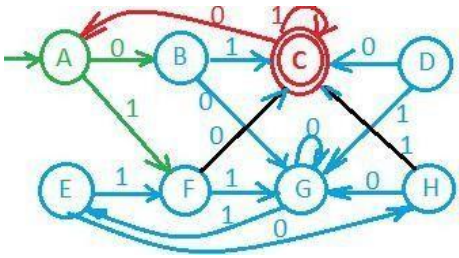


Figure 2.63: Deterministic finite automata.

From the given DFA we draw the dissimilar table as shown in figure. As compared to final state C all states are different (non-final states), hence X is marked.

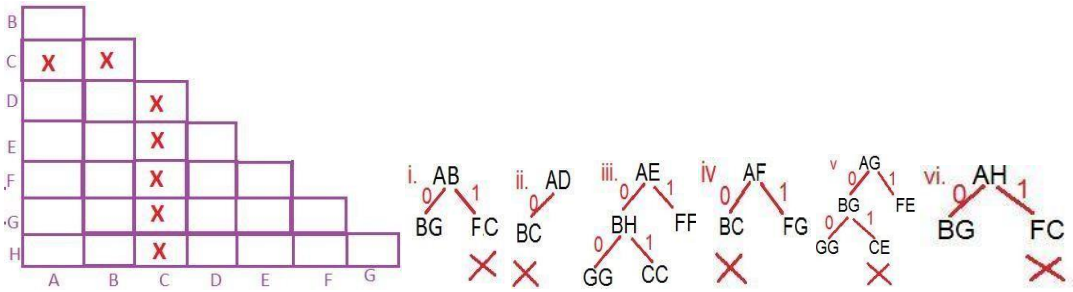


Figure 2.64: Differnce table with Shanprepat Tree.

The leaves node decides the root node. Any leaves node has X in the table, the root node is marked as X or else pair node repeats. Stop when the pair nodes start repeating, and decide the root node pair is similar.

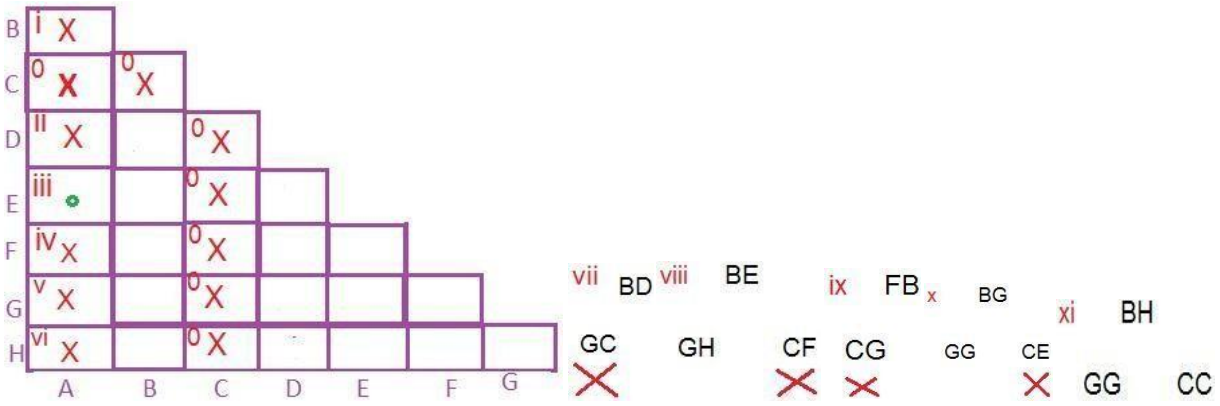


Figure 2.64: Differnce table with Shanprepat Tree.

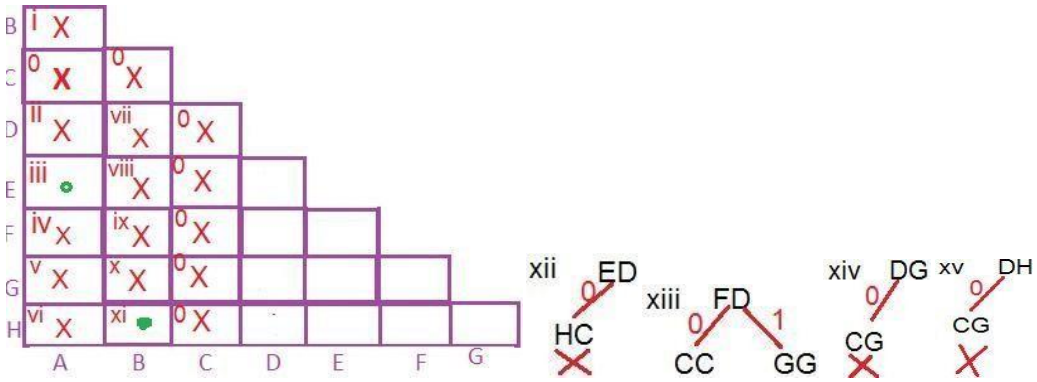


Figure 2.65: Differnce table with Shanprepat Tree.

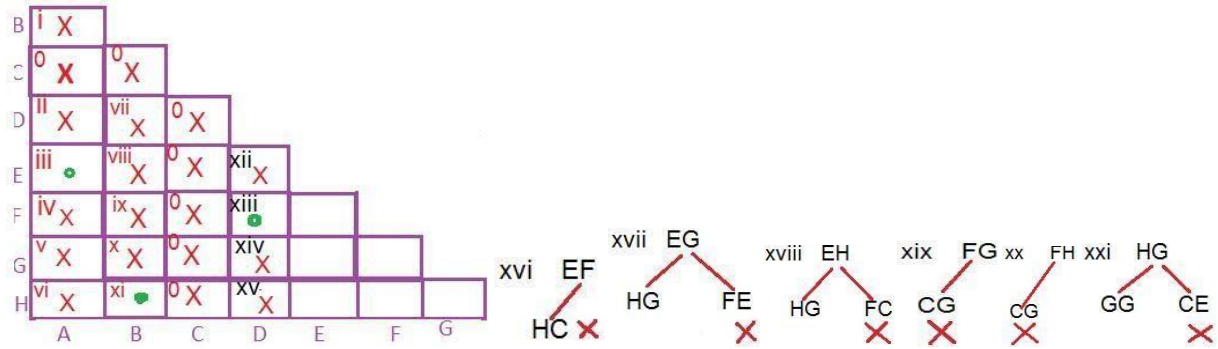


Figure 2.66: Differnce table with Shanprepat Tree.

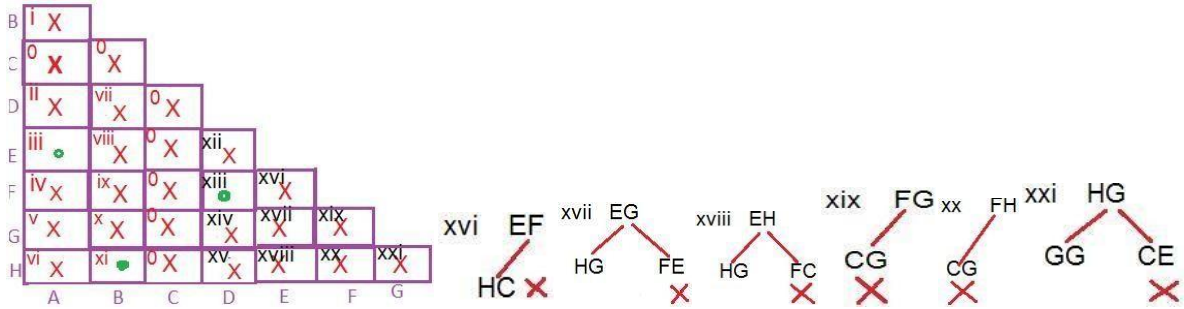


Figure 2.66: Differnce table with Shanprepat Tree.

From the table, we come to know that A=E, B=H and D=F. the resulting DFA is shown in figure..

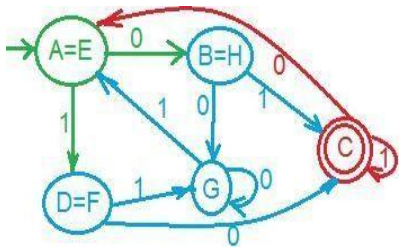


Figure 2.67: Minimized DFA of the figure 2.63

Example 2.14: Using table filling method, minimize the DFA shown in table and draw the transition diagram of the resulting DFA.

δ	0	1
→A	B	E
B	C	F
*C	D	H
D	E	H
E	F	I
*F	G	B
G	H	B
H	I	C
*I	A	E

The table shows the dissimilarity in the states of the given problem.

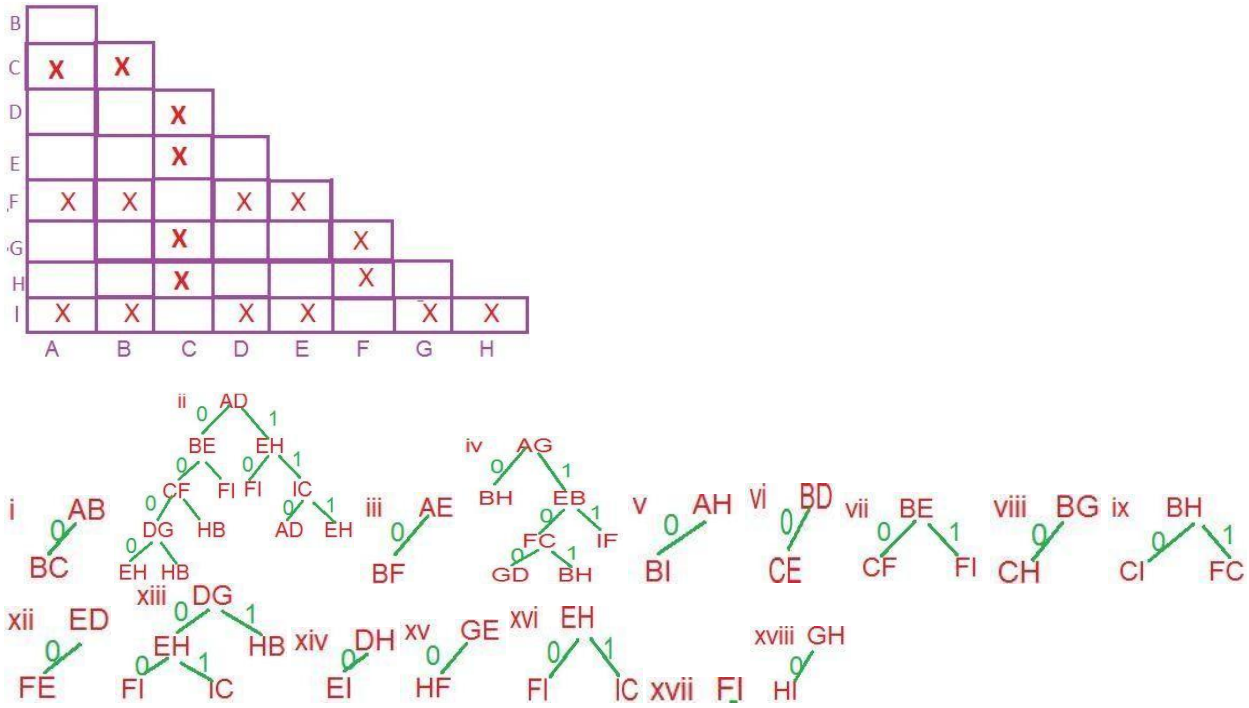


Figure 2.68: Differnce table with Shanprepat Tree.

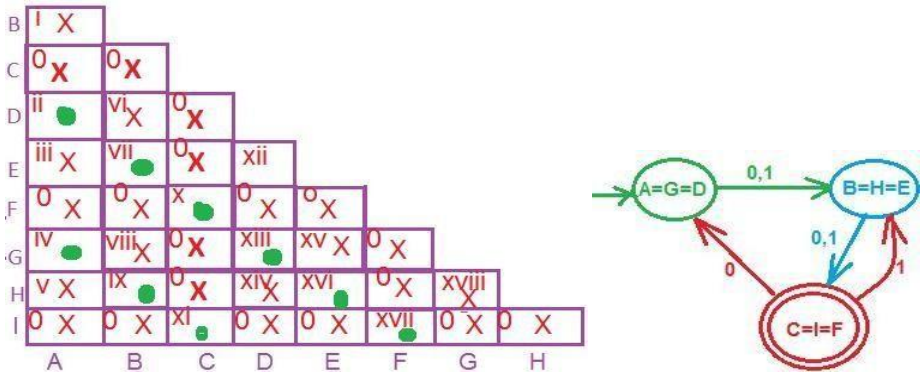


Figure 2.69: Differnce table with minimized DFA.

Module -III
Context-Free Grammars and Pushdown Automata (PDA)

1. Introduction to Rewrite Systems and Grammars

What is Rewrite System?

A rewrite system (or production system or rule based system) is a list of rules, and an algorithm for applying them. Each rule has a left-hand side and a right hand side.

$$X \rightarrow Y$$

(LHS) (RHS)

Examples of rewrite-system rules: $S \rightarrow aSb$, $aS \rightarrow \varepsilon$, $aSb \rightarrow bSabSa$

When a rewrite system R is invoked on some initial string w, it operates as follows:

simple-rewrite(R: rewrite system, w: initial string) =

1. Set working-string to w.
2. Until told by R to halt do:
   - **1.1** Match the LHS of some rule against some part of working-string.
   - **1.2** Replace the matched part of working-string with the RHS of the rule that was matched.
3. Return working-string.

If it returns some string s then R can derive s from w or there exists a derivation in R of s from w.

Examples:

1. A rule is simply a pair of strings where, if the string on the LHS matches, it is replaced by the string on the RHS.
2. The rule $axa \rightarrow aa$ squeeze out whatever comes between a pair of a's.
3. The rule $ab^*ab^*a \rightarrow aaa$ squeeze out b's between a's.

Rewrite systems can be used to define functions. We write rules that operate on an input string to produce the required output string. Rewrite systems can be used to define languages.The rewrite system that is used to define a language is called a grammar.

Grammars Define Languages

A grammar is a set of rules that are stated in terms of two alphabets:

- a terminal alphabet, Σ, that contains the symbols that make up the strings in L(G),
- a nonterminal alphabet, the elements of which will function as working symbols that will be used while the grammar is operating. These symbols will disappear by the time the grammar finishes its job and generates a string.
- A grammar has a unique start symbol, often called S.

A rewrite system formalism specifies:

- The form of the rules that are allowed.
- The algorithm by which they will be applied.
- How its rules will be applied?

Using a Grammar to Derive a String

Simple-rewrite (G, S) will generate the strings in L(G). The symbol ⇒ to indicate steps in a derivation.

Given: $S \rightarrow aS$------rule 1

$S \rightarrow \varepsilon$ ------rule 2

A derivation could begin with: $S \Rightarrow aSb \Rightarrow aaSbb \Rightarrow aabb$

Generating Many Strings

LHS of Multiple rules may match with the working string.

Given: $S \rightarrow aSb$----------rule 1

$S \rightarrow bSa$----------rule 2

$S \rightarrow \varepsilon$----------rule 3

Derivation so far: $S \Rightarrow aSb \Rightarrow aaSbb \Rightarrow$

Three are three choices at the next step:

$S \Rightarrow aSb \Rightarrow aaSbb \Rightarrow aaaSbbb$ (using rule 1),

$S \Rightarrow aSb \Rightarrow aaSbb \Rightarrow aabSabb$ (using rule 2),

$S \Rightarrow aSb \Rightarrow aaSbb \Rightarrow aabb$ (using rule 3).

One rule may match in more than one way.

Given: $S \rightarrow aTTb$----------rule 1

$T \rightarrow bTa$ ----------rule 2

$T \rightarrow \varepsilon$---------- rule 3

Derivation so far: $S \Rightarrow aTTb \Rightarrow$

Two choices at the next step:

$S \Rightarrow aTTb \Rightarrow abTaTb \Rightarrow$

$S \Rightarrow aTTb \Rightarrow aTbTab \Rightarrow$

When to Stop

Case 1: The working string no longer contains any nonterminal symbols (including, when it is ε). In this case, we say that the working string is generated by the grammar.

Example: $S \Rightarrow aSb \Rightarrow aaSbb \Rightarrow aabb$

Case 2: There are nonterminal symbols in the working string but none of them appears on the left-hand side of any rule in the grammar. In this case, we have a blocked or non-terminated derivation but no generated string.

Given: $S \rightarrow aSb$ ----- rule 1

$S \rightarrow bTa$ ----- rule 2

$S \rightarrow \varepsilon$ ----- rule 3

Derivation so far: $S \Rightarrow aSb \Rightarrow abTab \Rightarrow$

Case 3: It is possible that neither case 1 nor case 2 is achieved.

Given: $S \rightarrow Ba$---------- rule 1

$B \rightarrow bB$ ---------- rule 2

Then all derivations proceed as: $S \Rightarrow Ba \Rightarrow bBa \Rightarrow bbBa \Rightarrow bbbBa \Rightarrow bbbbBa \Rightarrow ...$

The grammar generates the language Ø.

2. Context –Free Grammar and Languages

Recall Regular Grammar which has a left-hand side that is a single nonterminal and have a right-hand side that is ε or a single terminal or a single terminal followed by a single nonterminal.

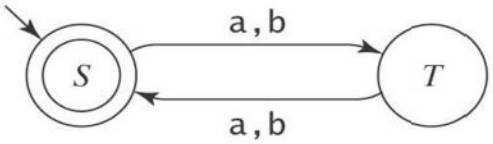
$$X \rightarrow Y$$

(NT) (ε or T or T NT)

Example: L = {w Î {a, b}* : |w| is even} RE = ((aa) (ab) (ba) (bb))*

G: $S \rightarrow \varepsilon$
$S \rightarrow aT$
$S \rightarrow bT$
$T \rightarrow aS$
$T \rightarrow bS$

M:



Context Free Grammars

$$X \rightarrow Y$$

(NT) (No restriction)

No restrictions on the form of the right hand sides. But require single non-terminal on left hand side.

Example: $S \rightarrow \varepsilon$, $S \rightarrow a$, $S \rightarrow T$, $S \rightarrow aSb$, $S \rightarrow aSbbT$ are allowed.

$ST \rightarrow aSb$, $a \rightarrow aSb$, $\varepsilon \rightarrow a$ are not allowed.

The name for these grammars "Context Free" makes sense because using these rules the decision to replace a nonterminal by some other sequence is made without looking at the context in which the nonterminals occurs.

Definition Context-Free Grammar

A context-free grammar G is a quadruple, (V, Σ, R, S), where:

- V is the rule alphabet, which contains nonterminals and terminals.
- Σ (the set of terminals) is a subset of V,
- R (the set of rules) is a finite subset of $(V - \Sigma) \times V^*$,
- S (the start symbol) is an element of $V - \Sigma$.

Given a grammar G, define $x \Rightarrow_G y$ to be the binary relation derives-in-one-step, defined so that $\forall x,y \in V^*$ ($x \Rightarrow_G y$ iff $x = \alpha A \beta$, $y = \alpha \gamma \beta$ and there exists a rule $A \rightarrow \gamma$ is in R_G)

Any sequence of the form $w_0 \Rightarrow_G w_1 \Rightarrow_G w_2 \Rightarrow_G \ldots \Rightarrow_G w_n$ is called a derivation in G. Let $\Rightarrow_G^*$ be the reflexive, transitive closure of $\Rightarrow_G$. We'll call $\Rightarrow_G^*$ the derive relation.

A derivation will halt whenever no rules on the left hand side matches against working-string. At every step, any rule that matches may be chosen.

Language generated by G, denoted L(G), is: $L(G) = \{w \in \Sigma^* : S \Rightarrow_G^* w\}$. A language L is context-free iff it is generated by some context-free grammar G. The context-free languages (or CFLs) are a proper superset of the regular languages.

Example: $L = A^nB^n = \{a^nb^n : n \geq 0\} = \{\varepsilon, ab, aabb, aaabbb, \ldots\}$

$G = \{\{S, a, b\}, \{a, b\}, R, S\}$, where: $R = \{ S \rightarrow aSb \ , \ S \rightarrow \varepsilon\}$

Example derivation in G: $S \Rightarrow aSb \Rightarrow aaSbb \Rightarrow aaaSbbb \Rightarrow aaabbb$ or $S \Rightarrow^* aaabbb$

Recursive Grammar Rules

A grammar is recursive iff it contains at least one recursive rule. A rule is recursive iff it is $X \rightarrow w_1Yw_2$, where: $Y \Rightarrow^* w_3Xw_4$ for some w_1, w_2, w_3, and w_4 in V^*. Expanding a non- terminal according to these rules can eventually lead to a string that includes the same non- terminal again.

Example1: $L = A^nB^n = \{a^nb^n : n \geq 0\}$ Let $G = (\{S, a, b\}, \{a, b\}, \{S \rightarrow a\ S\ b, S \rightarrow \varepsilon\}, S)$

Example 2: Regular grammar whose rules are $\{S \rightarrow a\ T, T \rightarrow a\ W, W \rightarrow a\ S, W \rightarrow a\ \}$

Example 3: The Balanced Parenthesis language

Bal = {w ∈ {),(}*: the parenthesis are balanced} = {ε , (), (()), ()(), (()()) .....}

G={{S,),(}, {),(},R,S} where R={ $S \rightarrow \varepsilon$ $S \rightarrow$ SS $S \rightarrow$ (S) }

Some example derivations in G:

$S \Rightarrow (S) \Rightarrow ()$

$S \Rightarrow (S) \Rightarrow (SS) \Rightarrow ((S)S) \Rightarrow (()\ S)) \Rightarrow (()\ (S)) \Rightarrow (()())$

So, S ⇒* () and S ⇒* (()())

Recursive rules make it possible for a finite grammar to generate an infinite set of strings.

Self-Embedding Grammar Rules

A grammar is self-embedding iff it contains at least one self-embedding rule. A rule in a grammar G is self-embedding iff it is of the form $X \rightarrow w_1 Y w_2$, where $Y \Rightarrow^* w_3 X w_4$ and both $w_1 w_3$ and $w_4 w_2$ are in Σ^+. No regular grammar can impose such a requirement on its strings.

Example: $S \rightarrow aSa$ is self-embedding

$S \rightarrow aS$ is recursive but not self- embedding

$S \rightarrow aT$

$T \rightarrow Sa$ is self-embedding

Example : PalEven = $\{ww^R : w \in \{a, b\}^*\}$= The L of even length palindrome of a's and b's.

L = {ε, aa, bb, aaaa, abba, baab, bbbb, ...}

G = {{S, a, b}, {a, b}, R, S}, where:

R = { $S \rightarrow aSa$ ---------- rule 1

$S \rightarrow bSb$ ---------- rule 2

$S \rightarrow \varepsilon$ ---------- rule 3 }.

Example derivation in G:

S ⇒ aSa ⇒ abSba ⇒ abba

Where Context-Free Grammars Get Their Power

If a grammar G is not self-embedding then L(G) is regular. If a language L has the property that every grammar that defines it is self-embedding, then L is not regular.

More flexible grammar-writing notations

a. Notation for writing practical context-free grammars. The symbol | should be read as "or". It allows two or more rules to be collapsed into one.

Example:

$S \rightarrow a\ S\ b$

$S \rightarrow b\ S\ a$ can be written as $S \rightarrow a\ S\ b \mid b\ S\ a \mid \varepsilon$

$S \rightarrow \varepsilon$

b. Allow a nonterminal symbol to be any sequence of characters surrounded by angle brackets.

Example1: BNF for a Java Fragment

<block> ::= {<stmt-list>} | {}

<stmt-list> ::= <stmt> | <stmt-list> <stmt>

<stmt> ::= <block> | while (<cond>) <stmt> |

if (<cond>) <stmt> |

do <stmt> while (<cond>); |

<assignment-stmt>; |

return | return <expression> |

<method-invocation>;

Example2: A CFG for C++ compound statements:

<compound stmt> → { <stmt list> }

<stmt list> → <stmt> <stmt list> | epsilon

<stmt> → <compound stmt>

<stmt> → if (<expr>) <stmt>

<stmt> → if (<expr>) <stmt> else <stmt>

<stmt> → while (<expr>) <stmt>

<stmt> → do <stmt> while (<expr>) ;

<stmt> → for (<stmt> <expr> ; <expr>) <stmt>

<stmt> → case <expr> : <stmt>

<stmt> → switch (<expr>) <stmt>

<stmt> → break ; | continue ;

<stmt> → return <expr> ; | goto <id> ;

Example3: A Fragment of English Grammar

Notational conventions used are

- Nonterminal = whose first symbol is an uppercase letter
- NP = derive noun phrase
- VP = derive verb phrase

S → NP VP

NP → the Nominal | a Nominal | Nominal | ProperNoun | NP PP

Nominal → N | Adjs N

N → cat | dogs | bear | girl | chocolate | rifle

ProperNoun → Chris | Fluffy

Adjs → Adj Adjs | Adj

Adj → young | older | smart

VP → V | V NP | VP PP

V → like | likes | thinks | shots | smells

PP → Prep NP

Prep → with

3. Designing Context-Free Grammars

If L has a property that every string in it has two regions & those regions must bear some relationship to each other, then the two regions must be generated in tandem. Otherwise, there is no way to enforce the necessary constraint.

Example 1: L = $\{a^n b^n c^m : n, m \geq 0\}$ = L = {ε, ab, c, abc, abcc, aabbc, …….}
The c^m portion of any string in L is completely independent of the $a^n b^n$ portion, so we should generate the two portions separately and concatenate them together.
G = ({S, A, C, a, b, c}, {a, b, c}, R, S} where:

R = { S → AC /* generate the two independent portions
A → aAb | ε /* generate the $a^n b^n$ portion, from the outside in
C → cC | ε } /* generate the c^m portion

Example derivation in G for string abcc:
S ⇒ AC ⇒ aAbC ⇒ abC ⇒ abcC ⇒ abccC ⇒ abcc

Example 2: L={ $a^i b^j c^k$: j=i+k, i ,k ≥ 0} on substituting j=i+k ⇒ L = $\{a^i b^i b^k c^k : i, k \geq 0\}$
L = {ε, abbc, aabbbbcc, abbbcc …….}
The $a^i b^i$ portion of any string in L is completely independent of the $b^k c^k$ portion, so we should generate the two portions separately and concatenate them together.
G = ({S, A, B, a, b, c}, {a, b, c}, R, S} where:

R = { S → AB /* generate the two independent portions
A → aAb | ε /* generate the $a^i b^i$ portion, from the outside in
B → bBc | ε } /* generate the $b^k c^k$ portion

Example derivation in G for string abbc:
S ⇒ AB ⇒ aAbB ⇒ abB ⇒ abbBc ⇒ abbc

Example 3: L={ $a^i b^j c^k$: i=j+k, j ,k ≥0} on substituting i=j+k ⇒ L = $\{a^k a^j b^j c^k : j, k \geq 0\}$
L = {ε, ac, ab, aabc, aaabcc, …….}
The $a^i b^i$ is the inner portion and $a^k c^k$ is the outer portion of any string in L.
G = ({S, A, a, b, c}, {a, b, c}, R, S} where:

R = { S → aSc | A /* generate the $a^k c^k$ outer portion
A → aAb | ε /* generate the $a^j b^j$ inner portion}

Example derivation in G for string aabc:
S ⇒ aSc ⇒ aAc ⇒ aaAbc ⇒ aabc

Example 4: L = $\{a^n w w^R b^n : w \in \{a, b\}^*\}$ = {ε, ab, aaab, abbb, aabbab, aabbbbab, ... }
The $a^n b^n$ is the inner portion and ww^R is the outer portion of any string in L.
G = {{S, A, a, b}, {a, b}, R, S}, where:

R = {S → aSb----------rule 1
S → A----------rule 2
A→ aAa ----------rule 3
A→ bAb ----------rule 4
A → ε----------rule 5 }.

Example derivation in G for string aabbab:
S ⇒ aSb ⇒ aAb ⇒ aaAab ⇒ aabAbab ⇒ aabbab

Example 5: Equal Numbers of a's and b's. = $\{w \in \{a, b\}^*: \#_a(w) = \#_b(w)\}$.

L = {ε, ab, ba, abba, aabb, baba, bbaa, …….}

G = {{S, a, b}, {a, b}, R, S}, where:

R = { $S \rightarrow aSb$ ----------rule 1

$S \rightarrow bSa$ ----------rule 2

$S \rightarrow SS$ ----------rule 3

$S \rightarrow \varepsilon$ ----------rule 4 }.

Example derivation in G for string abba:

$S \Rightarrow aSa \Rightarrow abSba \Rightarrow abba$

Example 6

$L = \{a^i b^j : 2i = 3j + 1\} = \{a^2b^1 , a^5b^3 , a^8b^5 \cdots\cdots\}$

G = {{S, a, b}, {a, b}, R, S}, where:

a^ib^j	$2i = 3j + 1$
a^2b^1	$2*2= 3*1 + 1 = 4$
a^5b^3	$2*5= 3*3 + 1 = 10$
a^8b^5	$2*8= 3*5 + 1 = 16$

R={ $S \rightarrow aaaSbb \mid aab$ }

Example derivation in G for string aaaaabbb:

$S \Rightarrow aaaSbb \Rightarrow aaaaabbb$

4. Simplifying Context-Free Grammars

Two algorithms used to simplify CFG

a. To find and remove unproductive variables using removeunproductive(G:CFG)
b. To find and remove unreachable variables using removeunreachable(G:CFG)

a. Removing Unproductive Nonterminals:

Removeunproductive (G: CFG) =

1. G′ = G.
2. Mark every nonterminal symbol in G′ as unproductive.
3. Mark every terminal symbol in G′ as productive.
4. Until one entire pass has been made without any new symbol being marked do:
   For each rule $X \rightarrow \alpha$ in R do:
   If every symbol in α has been marked as productive and X has not yet been marked as productive then:
   Mark X as productive.
5. Remove from G′ every unproductive symbol.
6. Remove from G′ every rule that contains an unproductive symbol.
7. Return G′.

Example: G = ({S, A, B, C, D, a, b}, {a, b}, R, S), where

R = { $S \rightarrow AB \mid AC$

$A \rightarrow aAb \mid \varepsilon$

$B \rightarrow bA$

$C \rightarrow bCa$

$D \rightarrow AB$ }

1) a and b terminal symbols are productive

2) A is productive(because $A \rightarrow aAb$)

3) B is productive(because $B \rightarrow bA$)

4) S & D are productive(because $S \rightarrow AB$ & $D \rightarrow AB$)

5) C is unproductive

On eliminating C from both LHS and RHS the rule set R′ obtained is

R′ = { $S \rightarrow AB$ $A \rightarrow aAb \mid \varepsilon$ $B \rightarrow bA$ $D \rightarrow AB$ }

b. Removing Unreachable Nonterminals

Removeunreachable (G: CFG) =

1. G′ = G.
2. Mark S as reachable.
3. Mark every other nonterminal symbol as unreachable.
4. Until one entire pass has been made without any new symbol being marked do:
   For each rule $X \rightarrow \alpha A \beta$ (where $A \in V - \Sigma$) in R do:
   If X has been marked as reachable and A has not then:
   Mark A as reachable.
5. Remove from G′ every unreachable symbol.
6. Remove from G′ every rule with an unreachable symbol on the left-hand side.
7. Return G′.

Example

G = ({S, A, B, C, D, a, b}, {a, b}, R, S), where

R′ = {$S \rightarrow AB$

$A \rightarrow aAb \mid \varepsilon$

$B \rightarrow bA$

$D \rightarrow AB$ }

S, A, B are reachable but D is not reachable, on eliminating D from both LHS and RHS the rule set R′′ is

R′′ = { $S \rightarrow AB$

$A \rightarrow aAb \mid \varepsilon$

$B \rightarrow bA$ }

5. Proving the Correctness of a Grammar

Given some language L and a grammar G, can we prove that G is correct (ie it generates exactly the strings in L)
To do so, we need to prove two things:

1. Prove that G generates only strings in L.
2. Prove that G generates all the strings in L.

6. Derivations and Parse Trees

Algorithms used for generation and recognition must be systematic. The expansion order is important for algorithms that work with CFG. To make it easier to describe such algorithms, we define two useful families of derivations.

a. A leftmost derivation is one in which, at each step, the leftmost nonterminal in the working string is chosen for expansion.
b. A rightmost derivation is one in which, at each step, the rightmost nonterminal in the working string is chosen for expansion.

Example 1 : S → AB | aaB A → a | Aa B → b
Left-most derivation for string aab is S ⇒ AB ⇒ AaB ⇒ aaB ⇒ aab
Right-most derivation for string aab is S ⇒ AB ⇒ Ab ⇒ Aab ⇒ aab

Example 2: S→iCtS | iCtSeS | x C→y
Left-most Derivation for string iytiytxex is S ⇒ iCtS ⇒ iytS ⇒ iytiCtSeS ⇒ iytiytSeS ⇒ iytiytxe ⇒ iytiytxex
Right-most Derivation for string iytiytxex is S ⇒ iCtSeS ⇒ iCtSex ⇒ iCtiCtSex ⇒ iCtiCtxex ⇒ iCtiytxex ⇒ iytiytxex

Example 3: A Fragment of English Grammar are
S → NP VP
NP → the Nominal | a Nominal | Nominal | ProperNoun | NP PP
Nominal → N | Adjs N
N → cat | dogs | bear | girl | chocolate | rifle
ProperNoun → Chris | Fluffy
Adjs → Adj Adjs | Adj
Adj → young | older | smart
VP → V | V NP | VP PP
V → like | likes | thinks | shots | smells
PP → Prep NP
Prep → with

Left-most Derivation for the string “the smart cat smells chocolate”

S ⇒ NP VP
⇒ the Nominal VP
⇒ the Adjs N VP
⇒ the Adj N VP
⇒ the smart N VP
⇒ the smart cat VP
⇒ the smart cat V NP
⇒ the smart cat smells NP
⇒ the smart cat smells Nominal
⇒ the smart cat smells N
⇒ the smart cat smells chocolate

Right-most Derivation for the string “the smart cat smells chocolate”

S ⇒ NP VP
⇒ NP V NP
⇒ NP V Nominal
⇒ NP V N
⇒ NP V chocolate
⇒ NP smells chocolate
⇒ the Nominal smells chocolate
⇒ the Adjs N smells chocolate
⇒ the Adjs cat smells chocolate
⇒ the Adj cat smells chocolate
⇒ the smart cat smells chocolate

Parse Trees

Regular grammar: in most applications, we just want to describe the set of strings in a language.
Context-free grammar: we also want to assign meanings to the strings in a language, for which we care about internal structure of the strings. Parse trees capture the essential grammatical structure of a string. A program that produces such trees is called a parser. A parse tree is an (ordered, rooted) tree that represents the syntactic structure of a string according to some formal grammar. In a parse tree, the interior nodes are labeled by non terminals of the grammar, while the leaf nodes are labeled by terminals of the grammar or ε.

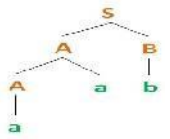
A parse tree, derived by a grammar G = (V, S, R, S), is a rooted, ordered tree in which:

1. Every leaf node is labeled with an element of $\sum \cup \{ \varepsilon \}$,
2. The root node is labeled S,
3. Every other node is labeled with some element of: $V - \sum$, and
4. If m is a nonleaf node labeled X and the children of m are labeled $x_1, x_2, \ldots, x_n$, then R contains the rule $X \rightarrow x_1, x_2, \ldots, x_n$.

Example 1: S → AB | aaB A → a | Aa B → b

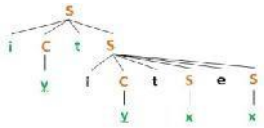
Left-most derivation for the string aab is S ⇒ AB ⇒ AaB ⇒ aaB ⇒ aab

Parse tree obtained is

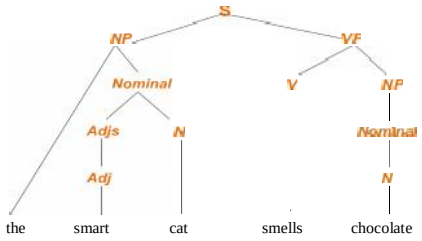


Example 2: S→iCtS | iCtSeS | x C→y

Left-most Derivation for string iytiytxex isS ⇒ iCtS ⇒ iytS ⇒ iytiCtSeS ⇒ iytiytSeS ⇒ iytiytxeS ⇒ iytiytxex



Example 3: Parse Tree -Structure in English for the string "the smart cat smells chocolate". It is clear from the tree that the sentence is not about cat smells or smart cat smells.



A parse tree may correspond to multiple derivations. From the parse tree, we cannot tell which of the following is used in derivation:

S ⇒ NP VP ⇒ the Nominal VP ⇒

S ⇒ NP VP ⇒ NP V NP ⇒

Parse trees capture the important structural facts about a derivation but throw away the details of the nonterminal expansion order. The order has no bearing on the structure we wish to assign to a string.

Generative Capacity

Because parse trees matter, it makes sense, given a grammar G, to distinguish between:

1. G's weak generative capacity, defined to be the set of strings, L(G), that G generates, and
2. G's strong generative capacity, defined to be the set of parse trees that G generates.

When we design grammar, it will be important that we consider both their weak and their strong generative capacities.

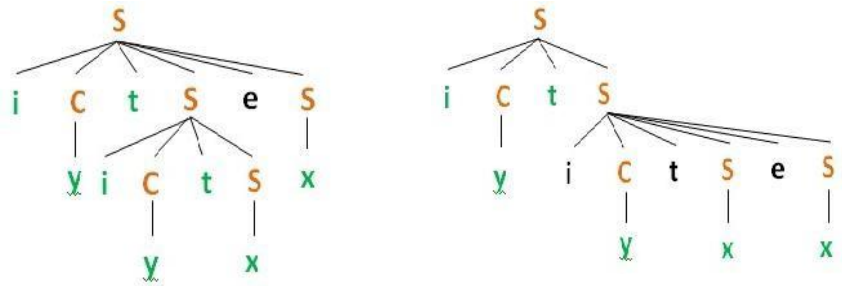
7. Ambiguity

Sometimes a grammar may produce more than one parse tree for some (or all) of the strings it generates. When this happens we say that the grammar is ambiguous. A grammar is ambiguous iff there is at least one string in L(G) for which G produces more than one parse tree.

Example 1: Bal={w ∈ {),(}*: the parenthesis are balanced}.
G={{S,),(}, {),(},R,S} where R={ S → ε S → SS S → (S) }
Left-most Derivation1 for the string (())() is S ⇒ S⇒(S)S ⇒ ((S))S ⇒ (())S ⇒ (())(S) ⇒ (())()
Left-most Derivation2 for the string (())() is S ⇒ SS ⇒SSS ⇒SS ⇒ (S)S ⇒ ((S))S ⇒ (())S ⇒ (())(S) ⇒ (())()

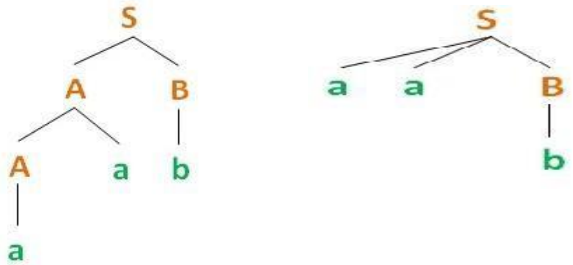
Since both the parse trees obtained for the same string (())() are different, the grammar is ambiguous.

Example 2: S→ iCtS | iCtSeS | x C→ y
Left-most Derivation for the string iytiytxex is S ⇒ iCtS ⇒ iytS ⇒ iytiCtSeS ⇒ iytiytSeS ⇒ iytiytxeS ⇒ iytiytxex
Right-most Derivation for the string iytiytxex is S ⇒ iCtSeS ⇒ iCtSex ⇒ iCtiCtSex ⇒iCtiCtxex ⇒ iCtiytxex ⇒ iytiytxex



Since both the parse trees obtained for the same string iytiytxex are different, the grammar is ambiguous.

Example 3: S → AB | aaB A → a | Aa B → b
Left-most derivation for string aab is S ⇒ AB ⇒ AaB ⇒ aaB ⇒ aab
Right-most derivation for string aab is S ⇒ aaB ⇒ aab



Since both the parse trees obtained for the same string aab are different, the grammar is ambiguous.

Why Is Ambiguity a Problem?

With regular languages, for most applications, we do not care about assigning internal structure to strings.

With context-free languages, we usually do care about internal structure because, given a string w, we want to assign meaning to w. It is generally difficult, if not impossible, to assign a unique meaning without a unique parse tree. So an ambiguous G, which fails to produce a unique parse tree is a problem.

Example : Arithmetic Expressions

G = (V, Σ, R, E), where

V = {+, *, (,), id, E},

Σ = {+, *, (,), id},

R = {E → E + E, E → E ∗ E, E → (E), E → id }

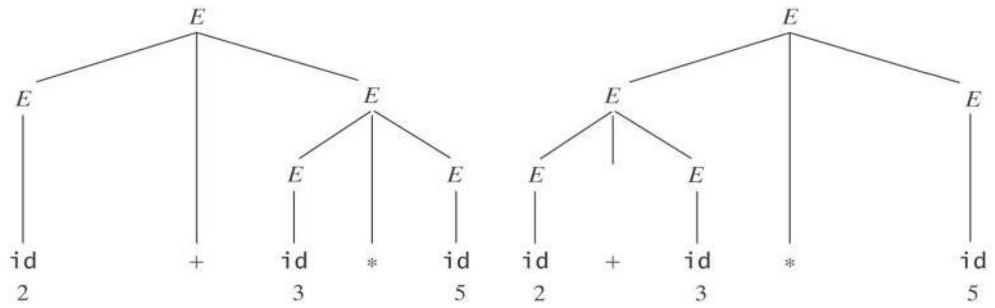
Consider string 2+3*5 written as id +id*id, left-most derivation for string id +id*id is

E ⇒ E*E ⇒ E+E*E ⇒ id+E*E ⇒ id+id*E ⇒ id+id*id.

Similarly the right-most derivation for string id +id*id is

E ⇒ E+E ⇒ E+E*E ⇒ E+E*id ⇒ E+id*id ⇒ id+id*id.

The parse trees obtained for both the derivations are:-



Should the evaluation of this expression return 17 or 25? Designers of practical languages must be careful that they create languages for which they can write unambiguous grammars.

Techniques for Reducing Ambiguity

No general purpose algorithm exists to test for ambiguity in a grammar or to remove it when it is found. But we can reduce ambiguity by eliminating

a. ε rules like S → ε
b. Rules with symmetric right-hand sides
    - A grammar is ambiguous if it is both left and right recursive.
    - Fix: remove right recursion
    - S → SS or E → E + E
c. Rule sets that lead to ambiguous attachment of optional postfixes.

a. Eliminating ε-Rules

Let G =(V, Σ, R, S) be a CFG. The following algorithm constructs a G′ such that L(G′) = L(G)-{ε} and G′ contains no ε rules:

removeEps(G: CFG) =

1. Let G′ = G.
2. Find the set N of nullable variables in G′.
3. Repeat until G′ contains no modifiable rules that haven't been processed:
   Given the rule $P \rightarrow \alpha Q \beta$, where $Q \in N$, add the rule $P \rightarrow \alpha\beta$ if it is not already present and if $\alpha\beta \neq \varepsilon$ and if $P \neq \alpha\beta$.
4. Delete from G′ all rules of the form $X \rightarrow \varepsilon$.
5. Return G′.

Nullable Variables & Modifiable Rules

A variable X is nullable iff either:

(1) there is a rule $X \rightarrow \varepsilon$, or

(2) there is a rule $X \rightarrow PQR\ldots$ and P, Q, R, … are all nullable.

So compute N, the set of nullable variables, as follows:

2.1. Set N to the set of variables that satisfy (1).

2.2. Until an entire pass is made without adding anything to N do

Evaluate all other variables with respect to (2).

If any variable satisfies (2) and is not in N, insert it.

A rule is modifiable iff it is of the form: $P \rightarrow \alpha Q \beta$, for some nullable Q.

Example: G = {{S, T, A, B, C, a, b, c}, {a, b, c}, R, S),

R = {$S \rightarrow aTa$ $T \rightarrow ABC$ $A \rightarrow aA \mid C$ $B \rightarrow Bb \mid C$ $C \rightarrow c \mid \varepsilon$ }

Applying removeEps

Step2: N = { C }

Step2.2 pass1: N = { A, B, C }

Step2.2 pass2: N = { A, B, C, T }

Step2.2 pass3: no new element found.

Step2: halts.

Step3: adds the following new rules to G′.

{ $S \rightarrow aa$

$T \rightarrow AB \mid BC \mid AC \mid A \mid B \mid C$

$A \rightarrow a$

$B \rightarrow b$ }

The rules obtained after eliminating ε-rules :

{ $S \rightarrow aTa \mid aa$

$T \rightarrow ABC \mid AB \mid BC \mid AC \mid A \mid B \mid C$

$A \rightarrow aA \mid C \mid a$

$B \rightarrow Bb \mid C \mid b$

$C \rightarrow c$ }

What If ε ∈ L?

Sometimes L(G) contains ε and it is important to retain it. To handle this case the algorithm used is

atmostoneEps(G: CFG) =

1. G′′ = removeEps(G).
2. If S_G is nullable then /* i. e., ε ∈ L(G)
   2.1 Create in G′′ a new start symbol S*.
   2.2 Add to $R_{G''}$ the two rules:S* → ε and S* → S_G.
3. Return G′′.

Example: Bal={w ∈ {),(}*: the parenthesis are balanced}.

R={ S → SS
S → (S)
S → ε }

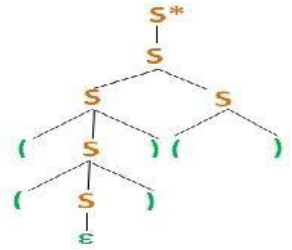
⟹

R={ S → SS
S → (S)
S → ( ) }

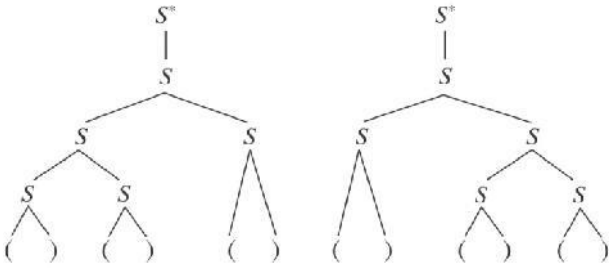
⟹

R={ S* → ε
S* → S
S → SS
S → (S)
S → ( )}

The new grammar built is better than the original one. The string (())() has only one parse tree.



But it is still ambiguous as the string ()()() has two parse trees?



Replace S → SS with one of:
S → S S_1 /* force branching to the left
S → S_1S /* force branching to the right

So we get: S* → ε | S
S → SS_1 /* force branching only to the left
S → S_1 /* add rule
S_1 → (S) | ()

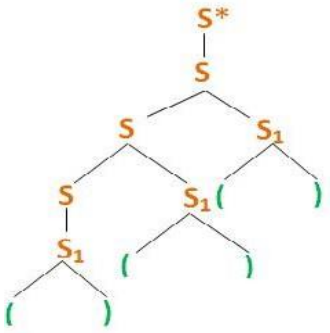
Unambiguous Grammar for Bal={w ∈ {),(}*: the parenthesis are balanced}.
G={{S,),(}, {),(},R,S} where

$S^* \rightarrow \varepsilon_n \mid S$

$S \rightarrow SS_1 \mid S_1$

$S_1 \rightarrow (S) \mid ()$

The parse tree obtained for the string ()()() is



Unambiguous Arithmetic Expressions

Grammar is ambiguous in two ways:

a. It fails to specify associatively.

Ex: there are two parses for the string id + id + id, corresponding to the bracketing (id + id) + id and id + (id + id)

b. It fails to define a precedence hierarchy for the operations + and *.

Ex: there are two parses for the string id + id * id, corresponding to the bracketing (id + id)* id and id + (id * id)

The unambiguous grammar for the arithmetic expression is:

$E \rightarrow E + T$

$E \rightarrow T$

$T \rightarrow T * F$

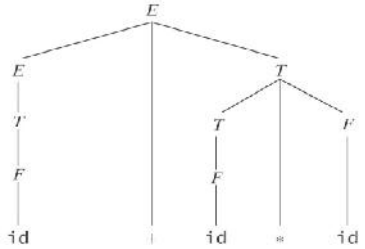
$T \rightarrow F$

$F \rightarrow (E)$

$F \rightarrow id$

For identical operators: forced branching to go in a single direction (to the left). For precedence Hierarchy: added the levels T (for term) and F (for factor)

The single parse tree obtained from the unambiguous grammar for the arithmetic expression is:



Proving that the grammar is Unambiguous

A grammar is unambiguous iff for all strings w, at every point in a leftmost derivation or rightmost derivation of w, only one rule in G can be applied.

$S^* \rightarrow \varepsilon$ ---(1)

$S^* \rightarrow S$ ---(2)

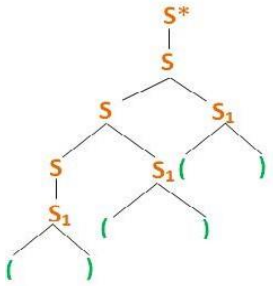
$S \rightarrow SS_1$ ---(3)

$S \rightarrow S_1$ ---(4)

$S_1 \rightarrow (S)$ ---(5)

$S_1 \rightarrow ()$ ---(6)

$S^* \Rightarrow S \Rightarrow SS_1 \Rightarrow SS_1S_1 \Rightarrow S_1S_1S_1 \Rightarrow () S_1S_1 \Rightarrow () () S_1 \Rightarrow () () ()$



Inherent Ambiguity

In many cases, for an ambiguous grammar G, it is possible to construct a new grammar G′ that generate L(G) with less or no ambiguity. However, not always. Some languages have the property that every grammar for them is ambiguous.We call such languages inherently ambiguous.

Example: $L = \{a^i b^j c^k : i, j, k \geq 0, i=j \text{ or } j=k\}$.

Every string in L has either (or both) the same number of a's and b's or the same number of b's and c's. $L = \{a^n b^n c^m : n, m \geq 0\} \cup \{a^n b^m c^m : n, m \geq 0\}$.

One grammar for L has the rules:

$S \rightarrow S_1 \mid S_2$

$S_1 \rightarrow S_1c \mid A$ /* Generate all strings in $\{a^n b^n c^m\}$.

$A \rightarrow aAb \mid \varepsilon$

$S_2 \rightarrow aS_2 \mid B$ /* Generate all strings in $\{a^n b^m c^m\}$.

$B \rightarrow bBc \mid \varepsilon$

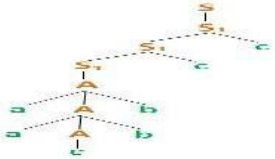
Consider the string $a^2b^2c^2$.

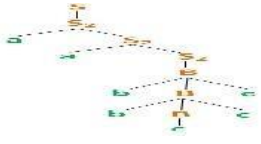
It has two distinct derivations, one through S_1 and the other through S_2

$S \Rightarrow S_1 \Rightarrow S_1c \Rightarrow S_1cc \Rightarrow Acc \Rightarrow aAbcc \Rightarrow aaAbbcc \Rightarrow aabbcc$

$S \Rightarrow S_2 \Rightarrow aS_2 \Rightarrow aaS_2 \Rightarrow aaB \Rightarrow aabBc \Rightarrow aabbBcc \Rightarrow aabbcc$

Given any grammar G that generates L, there is at least one string with two derivations in G.





Both of the following problems are undecidable:

- Given a context-free grammar G, is G ambiguous?
- Given a context-free language L, is L inherently ambiguous

8. Normal Forms

We have seen in some grammar where RHS is ε , it makes grammar harder to use. Lets see what happens if we carry the idea of getting rid of ε -productions a few steps farther. To make our tasks easier we define normal forms.

Normal Forms - When the grammar rules in G satisfy certain restrictions, then G is said to be in Normal Form.

- Normal Forms for queries & data can simplify database processing.
- Normal Forms for logical formulas can simplify automated reasoning in AI systems and in program verification system.
- It might be easier to build a parser if we could make some assumptions about the form of the grammar rules that the parser will use.

Normal Forms for Grammar

Among several normal forms, two of them are:-

- Chomsky Normal Form(CNF)
- Greibach Normal Form(GNF)

Chomsky Normal Form (CNF)

In CNF we have restrictions on the length of RHS and the nature of symbols on the RHS of the grammar rules.

A context-free grammar G = (V, Σ, R, S) is said to be in Chomsky Normal Form (CNF), iff every rule in R is of one of the following forms:

$X \rightarrow a$ where $a \in \Sigma$, or

$X \rightarrow BC$ where B and $C \in V\text{-}\Sigma$

Example: $S \rightarrow AB,\ A \rightarrow a, B \rightarrow b$

Every parse tree that is generated by a grammar in CNF has a branching factor of exactly 2 except at the branches that leads to the terminal nodes, where the branching factor is 1.

Using this property parser can exploit efficient data structure for storing and manipulating binary trees. Define straight forward decision procedure to determine whether w can be generated by a CNF grammar G. Easier to define other algorithms that manipulates grammars.

Greibach Normal Form (GNF)

GNF is a context free grammar G = (V, Σ, R, S), where all rules have one of the following forms: $X \rightarrow a\beta$ where $a \in \Sigma$ and $\beta \in (V\text{-}\Sigma)^*$

Example: $S \rightarrow aA \mid aAB,\ A \rightarrow a, B \rightarrow b$

In every derivation precisely one terminal is generated for each rule application. This property is useful to define a straight forward decision procedure to determine whether w can be generated by GNF grammar G. GNF grammars can be easily converted to PDA with no ε transitions.

Converting to Chomsky Normal Form

Apply some transformation to G such that the language generated by G is unchanged.

1. Rule Substitution.

Example: X→ aYc Y→ b Y → ZZ equivalent grammar constructed is X→ abc | aZZc

There exists 4-steps algorithm to convert a CFG G into a new grammar G_c such that: L(G) = $L(G_c) - \{\varepsilon\}$

convertChomsky(G:CFG) =

1. G' = removeEps(G:CFG) $S \to \varepsilon$
2. G'' = removeUnits(G':CFG) $A \to B$
3. G''' = removeMixed(G'':CFG) $A \to aB$
4. G'^{v} = removeLong(G''' :CFG) $S \to ABCD$

return G_c

Remove Epsilon using removeEps(G:CFG)

Find the set N of nullable variables in G.

X is nullable iff either $X \to \varepsilon$ or $(X \to A, A \to \varepsilon) : X \to \varepsilon$

Example1: G: $S \to aACa$

$A \to B \mid a$

$B \to C \mid c$

$C \to cC \mid \varepsilon$

Now, since C→ ε, C is nullable

since $B \to C$, B is nullable

since $A \to B$, A is nullable

Therefore N = { A,B,C}

removeEps returns G':

$S \to aACa \mid aAa \mid aCa \mid aa$

$A \to B \mid a$

$B \to C \mid c$

$C \to cC \mid c$

Remove Unit Productions using removeUnits(G:CFG)

Unit production is a rule whose right hand side consists of a single nonterminal symbol.

Ex: $A \to B$. Remove all unit production from G'.

Consider the remaining rules of G'.

G′: $S \to aACa \mid aAa \mid aCa \mid aa$

$A \to B \mid a$

$B \to C \mid c$

$C \to cC \mid c$

Remove $A \to B$ But $B \to C \mid c$, so Add $A \to C \mid c$

Remove $B \to C$ Add $B \to cC$ ($B \to c$, already there)

Remove $A \to C$ Add $A \to cC$ ($A \to c$, already there)

removeUnits returns G" :

$S \rightarrow aACa \mid aAa \mid aCa \mid aa$

$A \rightarrow cC \mid a \mid c$

$B \rightarrow cC \mid c$

$C \rightarrow cC \mid c$

Remove Mixed using removeMixed(G":CFG)

Mixed is a rule whose right hand side consists of combination of terminals or terminals with nonterminal symbol. Create a new nonterminal T_a for each terminal $a \in \Sigma$. For each T_a, add the rule $T_a \rightarrow a$

Consider the remaining rules of G" :

$S \rightarrow aACa \mid aAa \mid aCa \mid aa$

$A \rightarrow cC \mid a \mid c$

$B \rightarrow cC \mid c$

$C \rightarrow cC \mid c$

removeMixed returns G"' :

$S \rightarrow T_aACT_a \mid T_aAT_a \mid T_aCT_a \mid T_aT_a$

$A \rightarrow T_cC \mid a \mid c$

$B \rightarrow T_cC \mid c$

$C \rightarrow T_cC \mid c$

$T_a \rightarrow a$

$T_c \rightarrow c$

Remove Long using removeLong(G"' :CFG)

Long is a rule whose right hand side consists of more than two nonterminal symbol.

R: $A \rightarrow BCDE$ is replaced as: $A \rightarrow BM_2$

$M_2 \rightarrow CM_3$

$M_3 \rightarrow DE$

Consider the remaining rules of G"':

$S \rightarrow T_aACT_a \mid T_aAT_a \mid T_aCT_a$

Now, by applying removeLong we get :

$S \rightarrow T_aS_1$

$S_1 \rightarrow AS_2$

$S_2 \rightarrow CT_a$

$S \rightarrow T_aS_3$

$S_3 \rightarrow AT_a$

$S \rightarrow T_aS_2$

Now, by apply removeLong returns G^{iv} :

$S \rightarrow T_aS_1 \mid T_aS_3 \mid T_aS_2 \mid T_aT_a$

$S_1 \rightarrow AS_2$

$S_2 \rightarrow CT_a$

$S_3 \rightarrow AT_a$

$A \rightarrow T_cC \mid a \mid c$

$B \rightarrow T_cC \mid c$

$C \rightarrow T_cC \mid c$

$T_a \rightarrow a$

$T_c \rightarrow c$

Example 2: Apply the normalization algorithm to convert the grammar to CNF

G: $S \rightarrow aSa \mid B$

$B \rightarrow bbC \mid bb$

$C \rightarrow cC \mid \varepsilon$

removeEps(G:CFG) returns

G′: $S \rightarrow aSa \mid B$

$B \rightarrow bbC \mid bb$

$C \rightarrow cC \mid c$

removeUnits(G':CFG) returns

G′′ : $S \rightarrow aSa \mid bbC \mid bb$

$B \rightarrow bbC \mid bb$

$C \rightarrow cC \mid c$

removeMixed(G":CFG) returns

G''': $S \rightarrow T_aST_a \mid T_bT_bC \mid T_bT_b$

$B \rightarrow T_bT_bC \mid T_bT_b$

$C \rightarrow T_cC \mid c$

$T_a \rightarrow a$

$T_b \rightarrow b$

$T_c \rightarrow c$

removeLong(G''' :CFG) returns

G^{iv}: $S \rightarrow T_a S_1 \mid T_b S_2 \mid T_bT_b$

$S_1 \rightarrow S\ T_a$

$S_2 \rightarrow T_b\ C$

$B \rightarrow T_b\ S_2 \mid T_bT_b$

$C \rightarrow\ T_cC \mid c$

$T_a \rightarrow a$

$T_b \rightarrow b$

$T_c \rightarrow c$

Example 3: Apply the normalization algorithm to convert the grammar to CNF

G: $S \rightarrow ABC$
$A \rightarrow aC \mid D$
$B \rightarrow bB \mid \varepsilon \mid A$
$C \rightarrow Ac \mid \varepsilon \mid Cc$
$D \rightarrow aa$

removeEps(G:CFG) returns

G′: $S \rightarrow ABC \mid AC \mid AB \mid A$
$A \rightarrow aC \mid D \mid a$
$B \rightarrow bB \mid A \mid b$
$C \rightarrow Ac \mid Cc \mid c$
$D \rightarrow aa$

removeUnits(G':CFG) returns

G″ : $S \rightarrow ABC \mid AC \mid AB \mid aC \mid aa \mid a$
$A \rightarrow aC \mid aa \mid a$
$B \rightarrow bB \mid aC \mid aa \mid a \mid b$
$C \rightarrow Ac \mid Cc \mid c$
$D \rightarrow aa$

removeMixed(G":CFG) returns

G‴ : $S \rightarrow ABC \mid AC \mid AB \mid T_a C \mid T_a T_a \mid a$
$A \rightarrow T_a C \mid T_a T_a \mid a$
$B \rightarrow T_b B \mid T_a C \mid T_a T_a \mid a \mid b$
$C \rightarrow A T_c \mid C T_c \mid c$
$D \rightarrow T_a T_a$
$T_a \rightarrow a$
$T_b \rightarrow b$
$T_c \rightarrow c$

removeLong(G'" :CFG) returns

G′ᵛ: $S \rightarrow AS_1 \mid AC \mid AB \mid T_a C \mid T_a T_a \mid a$
$S_1 \rightarrow BC$
$A \rightarrow T_a C \mid T_a T_a \mid a$
$B \rightarrow T_b B \mid T_a C \mid T_a T_a \mid a \mid b$
$C \rightarrow A T_c \mid C T_c \mid c$
$D \rightarrow T_a T_a$
$T_a \rightarrow a$
$T_b \rightarrow b$
$T_c \rightarrow c$

9. Pushdown Automata

An acceptor for every context-free language. A pushdown automata , or PDA, is a finite state machine that has been augmented by a single stack.

Definition of a (NPDA) Pushdown Automaton

M = (K, S, G, Δ , s, A), where:

K is a finite set of states,
S is the input alphabet,
G is the stack alphabet,
s ∈ K is the initial state,
A ⊆ K is the set of accepting states, and
Δ is the transition relation.

Δ is the transition relation. It is a finite subset of

$$(K \times (\Sigma \cup \{\varepsilon\}) \times \Gamma^*) \times (K \times \Gamma^*)$$

- $(K \times (\Sigma \cup \{\varepsilon\}) \times \Gamma^*)$: state; input or ε; string of symbols to pop from top of stack
- $(K \times \Gamma^*)$: state; string of symbols to push on top of stack

Configuration

A configuration of PDA M is an element of K X S* X G*. Current state, Input that is still left to read and, Contents of its stack.

The initial configuration of a PDA M, on input w, is (s, w, ε).

will be written as cba

If $s_1s_2...s_n$ is pushed onto the stack: the value after the push is $s_1s_2...s_n$cba

Yields-in-one-step written $|-_M$ relates configuration$_1$ to
A computation by M is a finite sequence of configurations $C_0, C_1, C_2, \ldots\ldots\ldots\ldots C_n$ for some n ≥0 such that:

- C_0 is an initial configuration,
- C_n is of the form (q, ε, γ), for some q ∈ K and some string in G*, and
- $C_0 |-_M C_1 |-_M C_2 |-_M \ldots\ldots\ldots\ldots |-_M C_n$.

Nondeterminism
If M is in some configuration (q_1, s, γ) it s possible that:

- Δ contains exactly one transition that matches. In that case, M makes the specified move.
- Δ contains more than one transition that matches. In that case, M chooses one of them.
- Δ contains no transition that matches. In that case, the computation that led to that configuration halts.

Accepting
Let C be a computation of M on input w then C is an accepting configuration
iif C= (s, w, ε) $|-_M^*$ (q, ε, ε), for some q ∈ A.
A computation accepts only if it runs out of input w it is in an accepting state and the stack is empty.
C is a rejecting configuration iif C= (s, w, ε) $|-_M^*$ (q, w′, α),
where C is not an accepting computation and where M has no moves that it can makes from (q, w′, α). A computation can reject only if the criteria for accepting have not been met and there are no further moves that can be taken.
Let w be a string that is an element of S* . Then:

- M accepts w iif atleast one of its computations accepts.
- M rejects w iif all of its computations reject.

The language accepted by M, denoted L(M), is the set of all strings accepted by M. M rejectsa string w iff all paths reject it.
It is possible that, on input w, M neither accepts nor rejects. In that case, no path accepts and some path does not reject.

Transition
Transition ((q$_1$, c, γ1), $_2$, γ2)) ys that "If c ıc maes the input and g_1 matches the current top of the the s, tbe, the tran fron fr$_1$ to q_2 can bean be. Then, c will be removed from the input, 1 will γ1 wm be pom the stack, and 2 will bγ2 whed onto it. M cannot peek at the top of the stack vbe stack withg

- If c = ε, the tansition can be taken without consuming any input.
- If γ1 = ε, thetransition can be taken without checking the stack or popping anything. Note: it's not saying "the stack is empty".
- If γ2 = ε, nohing is pushed onto the stack when the transition is taken.

Example1: A PDA for Balanced Parentheses. Bal={w ∈ {),(}*: the parenthesis are balanced}

M = (K, S, G, Δ, s, A),

where:

K = {s} the states

S = {(,)} the input alphabet

Γ = {(} the stack alphabet

A = {s} the accepting state

Δ = {(s, ( )) ----------(1)

((s,), ( ) , (s,)) ---------(2) }

An Example of Accepting -- Input string = (())()

(S, (())(),) |- (S, ())(), ( ) |- (S,))()),(() |- (S,)(), () |- (S, (),) |- (S,), () |- (S, ε, ε)

The computation accepts the string (((())() as it runs out of input being in the accepting state S and stack empty.

Transition	State	Unread input	Stack
	S	(())()	
1	S	())()	(
1	S	))()	((
2	S	)()	(
2	S	()	
1	S	)	(
2	S		

Example1 of Rejecting -- Input string = (()))

(S, ())),ε) |- (S, ())),() |- (S,))),(() |- (S,)),() |- (S,), ε)

Transition	State	Unread input	Stack
	S	(()))	
1	S	()))	(
1	S	)))	((
2	S	))	(
2	S	)	

The computation has reached the final state S and stack is empty, but still the string is rejected because the input is not empty.

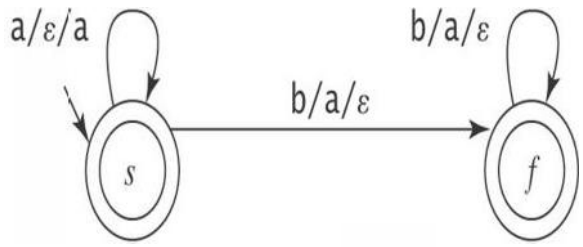
Example2 of Rejecting -- Input string = (())

Transition	State	Unread input	Stack
	S	(())	ε
1	S	())	(
1	S	))	((
2	S	)	((
2	S	ε	(

(S, (()),ε) |- (S, ()),(|- (S,( )),(() |- (S,)),((() |- (S,),(() |- (S, ε ,()
The computation has reached the final state S and runs out of input, but still the string is rejected because the stack is not empty.

Example 2: A PDA for $A^nB^n = \{a^nb^n: n \geq 0\}$
M = (K, S, G, Δ, s, A),
where:

K = {s, f} the states
S = {a, b} the input alphabetΓ
= {a} the stack alphabet
A = {s, f} the accepting state
Δ = { ((s, a, ε), (s, a)) -----(1)
((s, b, a), (f, ε)) -----(2)
((f, b, a), (f, ε)) } -----(3)



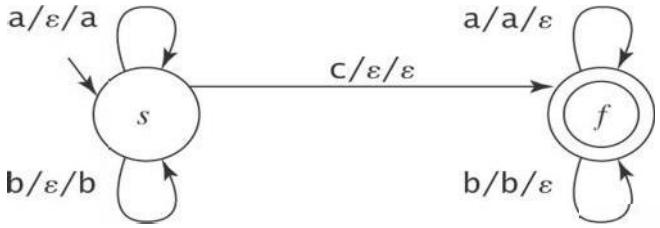
An Example of Accepting -- Input string = aabb
(f, aabb,) |- (f, abb, a) |- (f, bb, aa) |- (f, b, a) |- (f, ε, ε)

The computation has reached the final state f, the input string is consumed and the stack is empty. Hence the string aabb is accepted.

Example3: A PDA for $\{wcw^R: w \in \{a, b\}^*\}$
M = (K, S, G, Δ, s, A),
where:

K = {s, f} the states
S = {a, b, c} the input alphabet
Γ = {a, b} the stack alphabet
A = {f} the accepting state

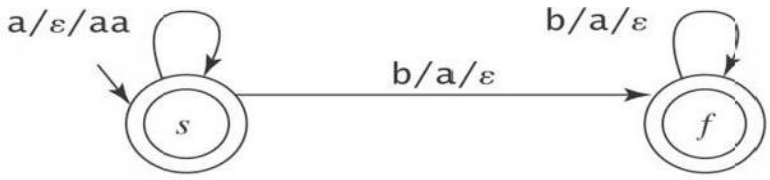
Δ = {((s, a, ε), (s, a)) -----(1)
((s, b, ε), (s, b)) -----(2)
((s, c, ε), (f, ε)) -----(3)
((f, a, a), (f, ε)) -----(4)
((f, b, b), (f, ε))} -----(5)



An Example of Accepting -- Input string = abcba
(s, abcba,) |- (s, bcba, a) |- (s, cba,ba) |- (f, ba, ba) |- (f, a, a) |- (f, ε, ε)
The computation has reached the final state f, the input string is co.consud and the stack is empty. Hence the string abcba is accepted.

Example 4: A PDA for A^nB^{2n} = {a^nb^{2n}: n ≥ 0}
M = (K, S, G, Δ, s, A),
where:
K = {s, f} the states
S = {a, b} the input alphabetΓ
= {a} the stack alphabet
A = {s, f} the accepting state

Δ = { ((s, a, ε), (s, aa)) ----- (1)
((s, b, a), (f, ε)) -----(2)
((f, b, a), (f, ε)) } -----(3)



An Example of Accepting -- Input string = aabbbb
(s, aabbbb,) |- (s, abbbb, aa) |- (s, bbbb,aaaa) |- (f, bbb, aaa) |- (f, bb, aa) |-(f, b, a) |- (f, ε, ε)

10. Deterministic and Nondeterministic PDAs

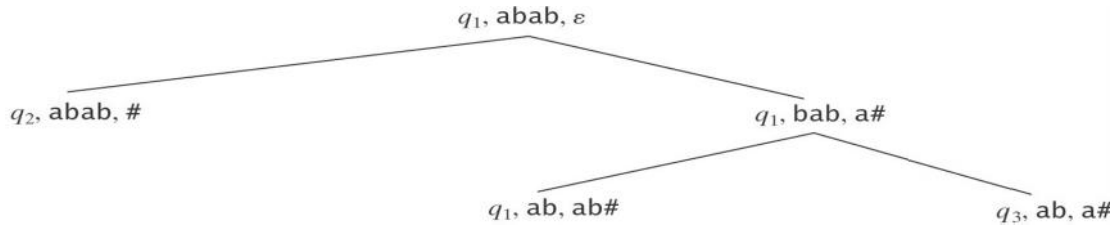
A PDA M is deterministic iff:

- Δ_M contains no pairs of transitions that compete with each other, and
- whenever M is in an accepting configuration it has no available moves.
- If q is an accepting state of M, then there is no transition ((q, e , e) ,(p, a)) for any p or a.

Unfortunately, unlike FSMs, there exist NDPDA s for which no equivalent DPDA exists.

Exploiting Nondeterministic

Previous examples are DPDA, where each machine followed only a single computational path. But many useful PDAs are not deterministic, where from a single configuration there exist multiple competing moves. As in FSMs, easiest way to envision the operation of a NDPDA M is as a tree.



Each node in the tree corresponds to a configuration of M and each path from the root to a leaf node may correspond to one computation that M might perform. The state, the stack and the remaining input can be different along different paths. As a result, it will not be possible to simulate all paths in parallel, the way we did for NDFSMs.

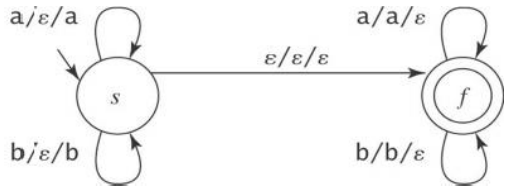
Example 1: PDA for PalEven ={ww^R: w ∈ {a, b}*}.

The L of even length palindrome of a's and b's. = {ε, aa, bb, aaaa, abba, baab, bbbb,.....}

M = (K, S, G, Δ, s, A),

where:

K = {s, f} the states
S = {a, b} the input alphabet
Γ = {a, b} the stack alphabet
A = {f} the accepting state
Δ = {((s, a, ε), (s, a)) -----(1)
((s, b, ε), (s, b)) -----(2)
((s, ε, ε), (f, ε)) -----(3)
((f, a, a), (f, ε)) -----(4)
((f, b, b), (f, ε))} -----(5)


a/ε/a
b/ε/b
ε/ε/ε
s
f
a/a/ε
b/b/ε

Example 2: PDA for {w ∈ {a, b}* : $\#_a(w) = \#_b(w)$}= Equal Numbers of a's and b's.

L = {ε, ab, ba, abba, aabb, baba, bbaa, …….}

M = (K, S, G, Δ, s, A),

where:

K = {s} the states

S = {a, b} the input alphabet

Γ = {a, b} the stack alphabet

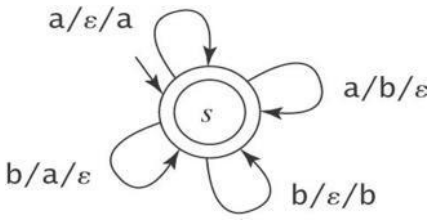
A = {s} the accepting state

Δ ={((s, a, ε), (s, a)) -----(1)

((s, b, ε), (s, b)) -----(2)

((s, a, b), (s, ε)) -----(3)

((s, b, a), (s, ε))} -----(4)

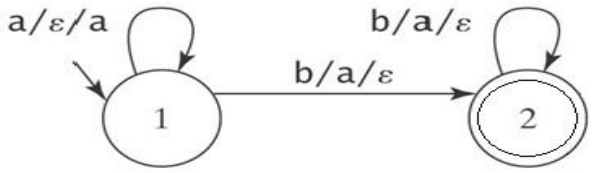


Example 3: The a Region and the b Region are Different. L = {$a^m b^n$: m ≠ n; m, n > 0}

It is hard to build a machine that looks for something negative, like ≠. But we can break L into two sublanguages: {$a^m b^n$: 0 < n < m} and {$a^m b^n$: 0 < m < n}

- If stack and input are empty, halt and reject
- If input is empty but stack is not (m > n) (accept)
- If stack is empty but input is not (m < n) (accept)

Start with the case where n = m

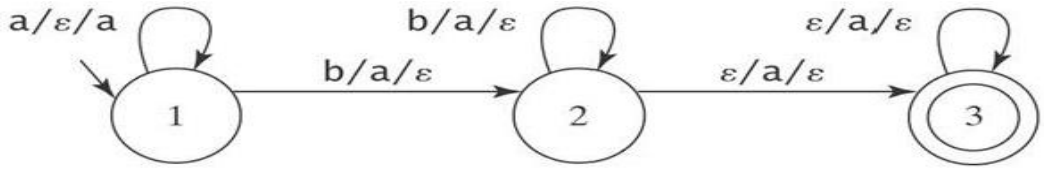


Δ = { ((1, a, ε), (1, a)) -----(1)

((1, b, a), (2, ε)) -----(2)

((2, b, a), (2, ε)) } ------(3)

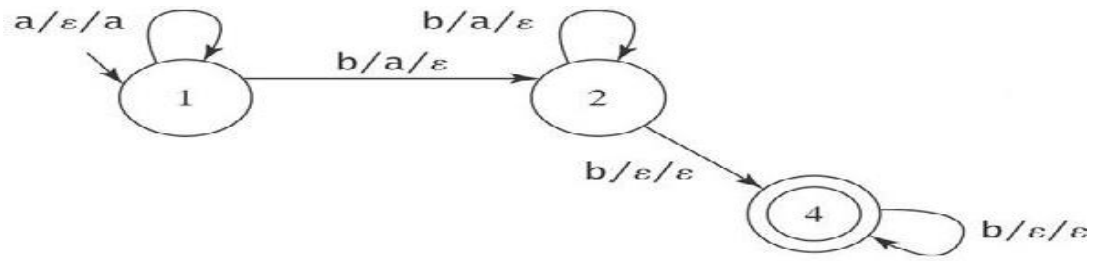
If input is empty but stack is not (m > n) (accept):



Δ = { ((1, a, ε), (1, a)) -----(1)
((1, b, a), (2, ε)) -----(2)
((2, b, a), (2, ε)) -----(3)
((2, ε, a), (3, ε)) -----(4)

((3, ε, a), (3, ε)) }------ (5)

If stack is empty but input is not (m < n) (accept):

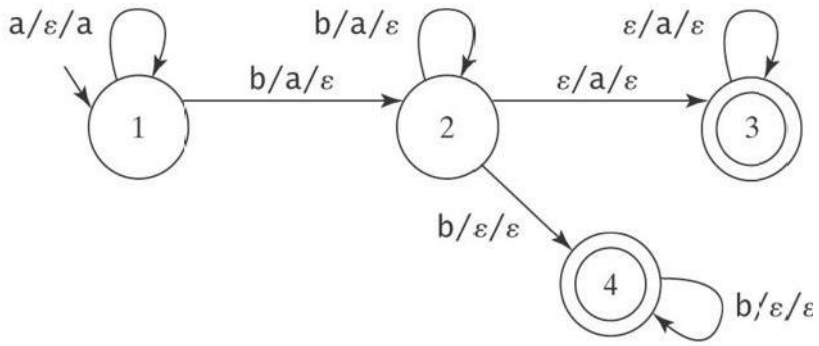


Δ = { ((1, a, ε), (1, a)) -----(1)
((1, b, a), (2, ε)) -----(2)
((2, b, a), (2, ε)) -----(3)
((2, b, ε), (4, ε)) -----(4)

((4, b, ε), (4, ε)) } ------(5)

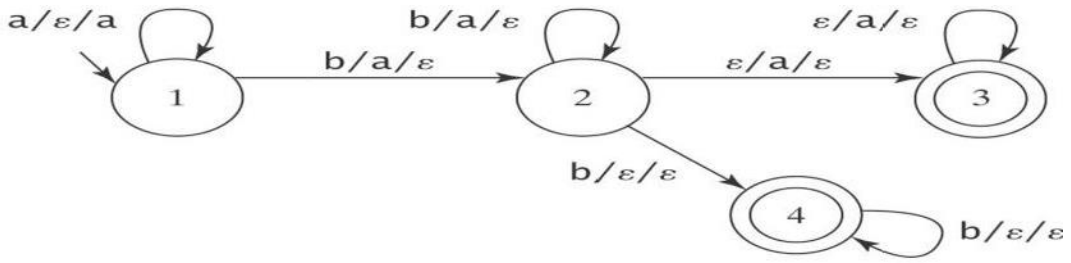
Putting all together the PDA obtained is
M = (K, S, G, Δ, s, A),
where:

K = {1,2,3,4} the states
S = {a, b} the input alphabet
Γ = {a} the stack alphabet
A = {3,4} the accepting state



Δ = { ((1, a, ε), (1, a)) -----(1)
((1, b, a), (2, ε)) -----(2)
((2, b, a), (2, ε)) -----(3)
((2, ε, a), (3, ε)) -----(4)
((3, ε, a), (3, ε)) } -----(5)
((2, b, ε), (4, ε)) -----(6)

((4, b, ε), (4, ε)) } -----(7)



Two problems with this M:

1. We have no way to specify that a move can be taken only if the stack is empty.
2. We have no way to specify that the input stream is empty.
3. As a result, in most of its moves in state 2, M will have a choice of three paths to take.

Techniques for Reducing Nondeterminism

We saw nondeterminism arising from two very specific circumstances:

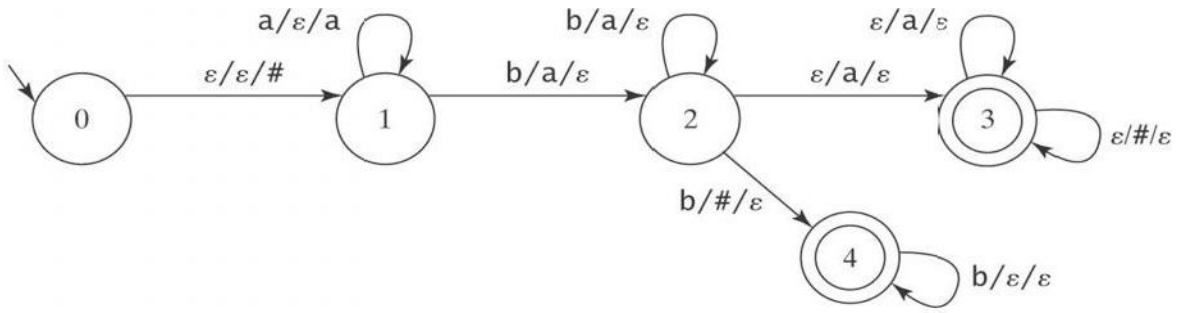
- A transition that should be taken only if the stack is empty competes against one or more moves that require a match of some string on the stack.
- A transition that should be taken only if the input stream is empty competes against one or more moves that require a match against a specific input character.

Case1: A transition that should be taken only if the stack is empty competes against one or more moves that require a match of some string on the stack.

Problem: Nondeterminism could be eliminated if it were possible to check for an empty stack.

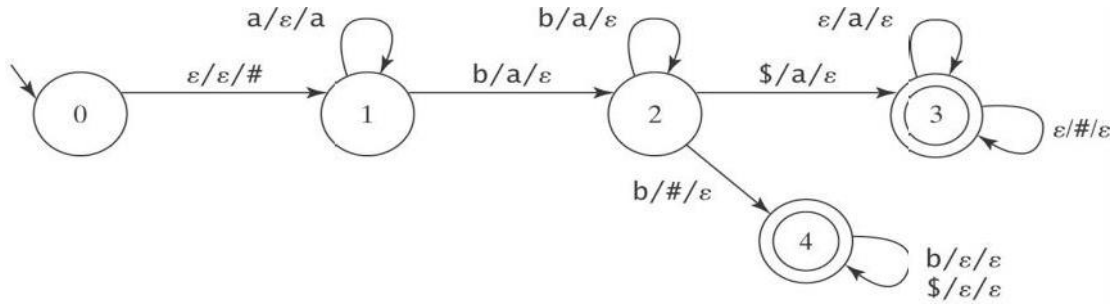
Solution: Using a special bottom-of-stack marker (#)

Before doing anything, push a special character onto the stack. The stack is then logically empty iff that special character (#) is at the top of the stack. Before M accepts a string, its stack must be completely empty, so the special character must be popped whenever M reaches an accepting state.



Now the transition back to state 2 no longer competes with the transition to state 4, which can only be taken when the # is the only symbol on the stack. The machine is still nondeterministic because the transition back to state 2 competes with the transition to state 3.

Case2: A transition that should be taken only if the input stream is empty competes against one or more moves that require a match against a specific input character.
Problem: Nondeterminism could be eliminated if it were possible to check for an empty input stream.
Solution: using a special end-of-string marker ($)
Adding an end-of-string marker to the language to be accepted is a powerful tool for reducing nondeterminism. Instead of building a machine to accept a language L, build one to accept L$, where $ is a special end-of-string marker.



Now the transition back to state 2 no longer competes with the transition to state 3, since the can be taken when the $ is read. The $ must be read on all the paths, not just the one where we need it.

11. Nondeterminism and Halting

Recall Computation C of a PDA M = (K, S, G, Δ, s, A) on a string w is an accepting

computation iif C= (s, w, ε) $|-_M^*$ (q, ε, ε), for some q ∈ A.
A computation C of M halts iff at least one of the following condition holds:

- C is an accepting computation, or
- C ends in a configuration from which there is no transition in Δ that can be taken.

M halts on w iff every computation of M on w halts. If M halts on w and does not accept, then we say that M rejects w.
For every CFL L, we've proven that there exists a PDA M such that L(M) = L.
Suppose that we would like to be able to:

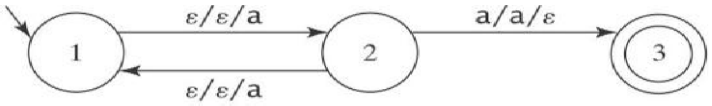
1. Examine a string and decide whether or not it is L.
2. Examine a string that is in L and create a parse tree for it.
3. Examine a string that is in L and create a parse tree for it in time that is linear in the length of the string.
4. Examine a string and decide whether or not it is in the complement of L.

For every regular language L, there exists a minimal deterministic FSM that accepts it. That minimal DFSM halts on all inputs, accepts all strings that are in L, and rejects all strings that are not in L.

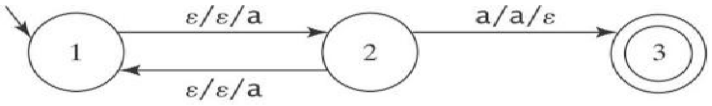
But the facts about CFGs and PDAs are different from the facts about RLs and FSMs.

1. There are context-free languages for which no deterministic PDA exists.
2. It is possible that a PDA may
   - not halt,
   - not ever finish reading its input.

However, for an arbitrary PDA M, there exists M′ that halts and L(M′) = L(M).
There exists no algorithm to minimize a PDA. It is undecidable whether a PDA is minimal.
Problem 2 : Let M be a PDA that accepts some language L. Then, on input w, if w ∈ L then M will halt and accept. But if w L then, while M will not accept w, it is possible that it will not reject it either.
Example1: Let S = {a} and consider M =



For L(M) = {a}. The computation (1, a, e) |- (2, a, a) |- (3, e, e) causes M to accept a.
Example2: Consider M =



For L(M) = {aa} or on any other input except a:
(1, aa, e) |- (2, aa, a) |-(1, aa, aa) |- (2, aa, aaa) |- (1, aa, aaaa) |- (2, aa, aaaaa) |- ……..
M will never halt because of one path never ends and none of the paths accepts.
The same problem with NDFSMs had a choice of two solutions.

- Converting NDFSM to and equivalent DFSM using ndfsmtodfsm algorithm.
- Simulating NDFSM using ndfsmsimulate.

Neither of these approaches work on PDAs. There may not even be an equivalent deterministic PDA.
Solution fall into two classes:

- Formal ones that do not restrict the class of the language that are being considered- converting grammar into normal forms like Chomsky or Greibach normal form.
- Practical ones that work only on a subclass of the CFLs- use grammars in natural forms.

12. Alternative Equivalent Definitions of a PDA

PDA M = (K, S, G, Δ , s, A):

1. Allow M to pop and to push any string in G*.
2. M may pop only a single symbol but it may push any number of them.
3. M may pop and push only a single symbol.

M accepts its input w only if , when it finishes reading w, it is in an accepting state and its stack is empty.

There are two alternatives to this:

1. PDA by Final state: Accept if, when the input has been consumed, M lands in an accepting state, regardless of the contents of the stack.
2. PDA by Empty stack: Accept if, when the input has been consumed, the stack is empty, regardless of the state M is in.

All of these definitions are equivalent in the sense that, if some language L is accepted by a PDA using one definition, it can be accepted by some PDA using each of the other definition. For example:- If some language L is accepted by a PDA by Final state then it can be accepted by PDA by Final state and empty stack. If some language L is accepted by a PDA by Final state and empty stack then can be accepted by PDA by Final state.

We can prove by showing algorithms that transform a PDA of one sort into and equivalent PDA of the other sort.

Equivalence

1. Given a PDA M that accepts by accepting state and empty stack, construct a new PDA M′ that accepts by accepting state alone, where L(M′) = L(M).
2. Given a PDA M that accepts by accepting state alone, construct a new PDA M′ that accepts by accepting state and empty stack, where L(M′) = L(M).

Hence we can prove that M′ and M accept the same strings.

1. Accepting by Final state Alone

Define a PDA M = (K, S, G, Δ , s, A). Accepts when the input has been consumed, M lands in an accepting state, regardless of the contents of the stack. M accepts if C= (s, w, ε) $|-_M^*$ (q, ε, g), for some q ∈ A.

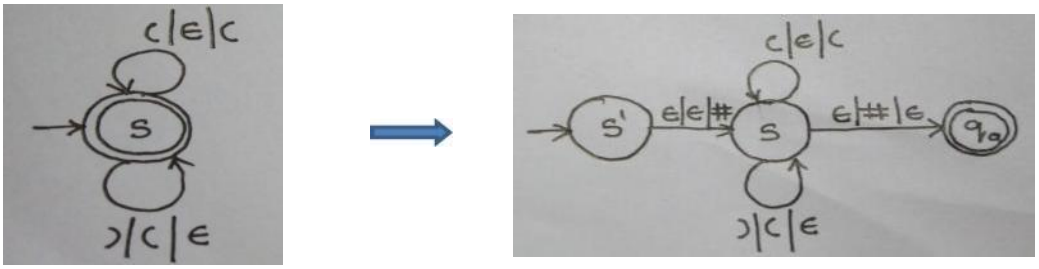
M′ will have a single accepting state q_a. The only way for M′ to get to q_a will be to land in an accepting state of M when the stack is logically empty. Since there is no way to check that the stack is empty, M′ will begin by pushing a bottom-of-stack marker #, onto the stack.

Whenever # is the top symbol of the stack, then stack is logically empty.

The construction proceeds as follows:

1. Initially, let M′ = M.
2. Create a new start state s′.
   Add the transition ((s′, ε , ε),(s, #)),
3. For each accepting state a in M do:
   Add the transition ((a, ε ,#),(q_a, ε)),
4. Make q_a the only accepting state in M′

Example:



It is easy to see that M′ lands in its accepting state(q_a) iff M lands in some accepting state with an empty stack. Thus M′ and M accept the same strings.

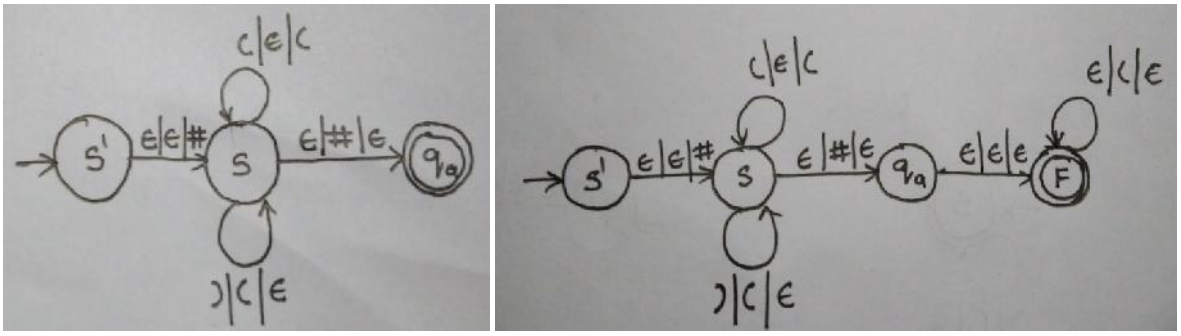
2. Accepting by Final state and Empty stack

The construction proceeds as follows:

1. Initially, let M′ = M.
2. Create a new accepting state F
3. From each accepting state a in M do:
   Add the transition ((a, ε , ε),(F, ε)),
4. Make F the only accepting state in M′
5. For every element g of Γ,
   Add the transition to M′ ((F, ε, g), (F, ε)).

In other words, iff M accepts, go to the only accepting state of M′ and clear the stack. Thus M′ will accept by accepting state and empty stack iff M accepts by accepting state.

Example:-



Thus M′ and M accept the same strings.

13. Alternatives that are not equivalent to the PDA

We defined a PDA to be a finite state machine to which we add a single stack.

Two variants of that definition, each of which turns out to define a more powerful class of a machine.

1. First variant: add a first-in, first-out (FIFO) queue in place of a stack. Such machines are called tag systems or Post machines.
2. Second variant: add two stacks instead of one. The resulting machines are equivalent in computational power to Turing Machines.

Sl.No	Sample Questions
1.	Define context free grammars and languages.
2.	Show a context-free grammar for each of the following languages L: a) BalDelim = {w : where w is a string of delimeters: (,), [,], {, }, that are properly balanced}. b) $\{a^ib^j : 2i = 3j + 1\}$. c) $\{a^ib^j : 2i \neq 3j + 1\}$. d) $\{a^ib^jc^k : i, j, k \geq 0$ and $(i \neq j$ or $j \neq k)\}$.
3.	Define CFG. Design CFG for the language $L=\{ a^nb^m : n \neq m\}$
4.	Apply the simplification algorithm to simplify the given grammar S → AB|AC A → aAb | ε B → bA C →bCa D → AB
5.	Prove the correctness of the grammar for the language: $L=\{w \in \{a, b\}^* : \#_a(w) = \#_b(w)\}$.
6.	Define leftmost derivation and rightmost derivation. Given the following CFG. E→ E + T|T T→ T*F|F F→(E)|a|b|c Draw parse tree for the following sentences and also derive the leftmost and rightmost derivations i) (a+b)*c ii) (a) + b*c
7.	Consider the following grammar G: S → 0S1 | SS | 10 Show a parse tree produced by G for each of the following strings: a) 010110 b) 00101101
8.	Define ambiguous and explain inherently ambiguous grammars.
9.	Prove whether the given grammar is ambiguous grammar or not. E→ E + E E→ E*E|a|b|c
10.	Prove that the following CFG is ambiguous S→iCtS|iCtSeS|x C→y for the sting iytiytxex
11.	Define Chomsky normal form. Apply the normalization algorithm to convert the grammar to Chomsky normal form. a. S → aSa S → B B → bbC B → bb C → ε C → cC b. S → ABC A → aC | D B → bB | ε | A C → Ac | ε | Cc D → aa
12.	Define Push down automata (NPDA). Design a NPDA for the CFG given in Question (2).
13.	Design a PDA for the given language.L$, where $L = \{w \in \{a, b\}^* : \#_a(w) = \#_b(w)\}$.
14.	Design a PDA for the language: $L=\{ a^ib^jc^k : i+j=k, i>=0, j>=0\}$
15.	Design a PDA for the language $L=\{ a^nb^{2n} : n>=1\}$
16.	Design a PDA for the language: $L=\{ a^ib^jc^k : i+k=j, i>=0, k>=0\}$
17.	Design a PDA for the language: $L=\{ a^ib^jc^k : k+j=i, k>=0, j>=0\}$

Module-4

Context-Free and Non-Context-Free Languages

- The language $A^nB^n = \{a^nb^n \mid n \geq 0\}$ is context-free.
- The language $A^nB^nC^n = \{ a^nb^nc^n \mid n \geq 0\}$ is not context free because a PDA's stack cannot count all three of the letter regions and compare them.

Where Do the Context-Free Languages Fit in the Big Picture?

THEOREM: The Context-Free Languages Properly Contain the Regular Languages.

Theorem: The regular languages are a proper subset of the context-free languages.

Proof: We first show that every regular language is context-free. We then show that there exists at least one context-free language that is not regular.

Every regular language is context-free : We show this by construction.

- If L is regular then it is accepted by some DFSM M = (K, **Σ**, δ, s, A).
- From M we construct a PDA
  M' = (K',**Σ', Γ',Δ',s',A'**) to accept *L*. where Δ' is constructed as follows:
  For every transition (q_i,c,q_j)in δ, add to Δ' the transition $((q_i,c,\varepsilon),(q_i,\varepsilon))$,So $L(M)=L(M')$.

So, the set of regular languages is a subset of the CFL.

There exists at least one context-free language that is not regular : The regular languages are a *proper* subset the context-free languages because there exists at least one language $\mathbf{a^nb^n}$ that is context –free but not regular.

Theorem: There is a countably infinite number of context-free languages.

Proof:

Every context-free language is generated by some context-free grammar G = (V,Σ,R,S).

There cannot be more CFLs than CFGs. So there are at most a countably infinite number of context-free languages. There is not a one-to-one relationship between CFLs and CFGs, since there are an infinite number of grammars that generate any given language. But we know that, every regular language is context free and there is a countably infinite number of regular languages.

So there is at least and at most a countably infinite number of CFLs.

Showing That a Language is Context-Free

Two techniques that can be used to show that language L is context-free:

•Exhibit a context-free grammar for it.

•Exhibit a (possibly nondeterministic) PDA for it.

Theorem: The length of the yield of any tree T with height h and branching factor b is <= b^h.

Proof:

If h is 1, then a single rule applies. So the longest yield is of length less than or equal to b. Assume the claim is true for h=n. We show that it is true for h=n+1.

Consider any tree with h=n+1. It consists of a root, and some number of subtrees, each of height <=n. By the induction hypothesis, the length of the yield of each of those subtrees is <= b^n. So the length of the yield must be <=b.(b^n)=b^{n+1}=b^h.

The Pumping Theorem for Context-Free languages

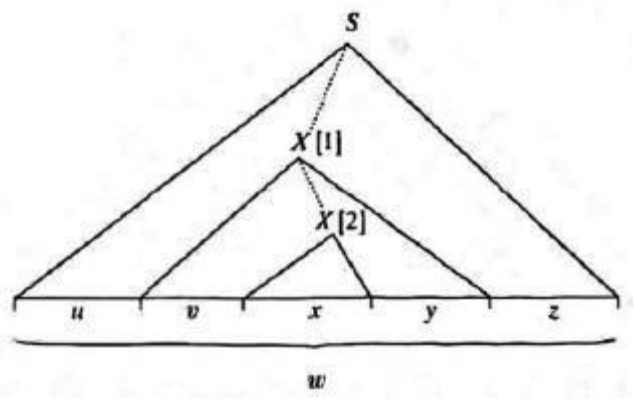
Statement: If L is CFL, then: ∃k≥1 (∀strings *w*∈L, where |w|≥k (∃u,v,x,y,z (*w*=uvxyz, vy≠Ɛ, |vxy| ≤ k and ∀q ≥ 0 (uv^qxy^qz is in L))))

Proof: If L is context-free, then there exists a CFG G=(V,Σ,R,S) with n nonterminal symbols and branching factor b.

Let *k* be b^{n+1}.

Any string that can be generated by G and whose parse tree contains no paths with repeated nonterminals must have length less than or equal to b^n. Assuming that b≥2, it must be the case that b^{n+1} > b^n. So let w be any string in L(G) where |w|≥k.

Let T be any smallest parse tree for w. T must have height at least n+1. Choose some path in T of length at least n + 1. Let X be the bottom-most repeated non terminal along that path. Then w can be rewritten as uvxyz as shown in below tree,



$S \Rightarrow^* uXz \Rightarrow^* uvXyz \Rightarrow^* uvxyz.$

The tree rooted at [1] has height at most n+1.Thus its yield, *vxy*, has length less than or equal to b^{n+1},which is k. Further, vy≠Ɛ .Since if *vy* were Ɛ then there would be a smal er parse tree for wand we choose T so that h at wasn't so.

Finally, v and y can be pumped: *uxz* must be in L because rule2 could have been used immediately

at[1]. And, for any $q \geq 1$, $uv^q xy^q z$ must be in L because rule1 could have been used q times before finally using rule2.

Application of pumping lemma (Proving Language is Not Context Free)

Ex1: Prove that the Language L = {$a^n b^n c^n$| n>=0} is Not Context-Free.

Solution: If L is CFL then there would exist some k such that any string *w*, where |w|>=k must satisfy the conditions of the theorem.

Let $w = a^k b^k c^k$, where 'k' is the constant from the Pumping lemma theorem. For *w* to satisfy the conditions of the Pumping Theorem there must be some u,v,x,y and z, such that w=uvxyz, vy≠Ɛ, |vxy|≤k and $\forall q \geq 0$, $uv^q xy^q z$ is in L.

w=aaa…aaabbb…bbbccc…ccc, select v and y as follows:

w=aaa…a**aabbb…bbbcc**c…ccc
v y

Let q=2, then

V^2 y^2
w=aaa…aaabbaabb b..bbccccc…ccc

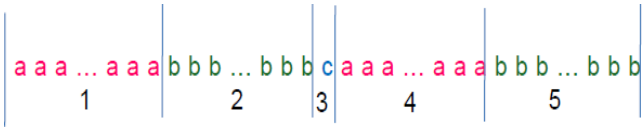
The resulting string will have letters out of order and thus not in L.

So L is not context-free.

Ex 2: Prove that the Language L= {WcW: w∈{a,b}*}is Not Context-Free.

For w to satisfy the conditions of the Pumping Theorem there must be some u,v,x,y,and z, such that w = uvxyz , vy≠Ɛ, |vxy| ≤ k and ∀q≥0, uv^qxy^qz is in L. We show that no such u,v,x,y and z exist.

Imagine w divided into five regions as follows:



Call the **part before the c the leftside and the part after the c the right side.** We consider all the cases for where v and y could fall and show that in none of them are all the condition so f the theorem met:

• If either v or y overlaps region 3, set q to 0. The resulting string will no longer contain a c and so is not in WcW.

• If both v and y occur before region 3 or they both occur after region 3, then set q to 2. One side will be longer than the other and so the resulting string is not in WcW.

• If either v or y overlaps region1 ,then set q to 2. In order to make the right side match. Something would have to be pumped into region 4. But any v,y pair that did that would violate the requirement that lvxyl ≤ k.

• If either v or y overlaps region2, then set q to 2. In order to make the right side match, something would have to be pumped into region 5. But any v,y pair that did that would violate the requirement that lvxyl ≤ k.

• There is no way to divide w into uvxyz such that all the conditions of the Pumping Theorem are met . So WcW is not context-free.

Some Important Closure Properties of Context-Free Languages

Theorem: The context- free languages are closed under Union, Concatenation, Kleene star, Reverse, and Letter substitution.

(1) The context-free languages are closed under union:

• If L1 and L2 are context free languages then there exists a context-free grammar $G_1 = (V_1, \Sigma_1, R_1, S_1)$ and $G_2 = (V_2, \Sigma_2, R_2, S_2)$ such that $L_1 = L(G_1)$ and $L_2 = L(G_2)$.

• We will build a new grammar G such that $L(G)=L(G_1)UL(G_2)$. G will contain all the rules of both G_1 and G_2.

• We add to G a new start symbol S and two new rules. $S\rightarrow S_1$ and $S\rightarrow S_2$. The two new rules allow G to generate a string iff at least one of G_1or G_2 generates it.

So, G = (V_1 U V_2 U {S}, Σ_1 U Σ_2, R_1 U R_2 U {S→ S_1,S→S_2}, S)

(2) The context-free languages are closed under concatenation

• If L_1 and L_2 are context free languages then there exist context-free grammar $G_1= (V_1,\Sigma_1,R_1,S_1)$ and $G_2=(V_2,\Sigma_2,R_2,S_2)$ such that $L_1= L(G_1)$ and $L_2= L(G_2)$.

• We will build a new grammar G such that $L(G) = L(G_1)L(G_2)$.

• G will contain all the rules of both $G_1$and G_2.

• We add to G a new start symbol S and one new rule. $S\rightarrow S_1S_2$

So $G= (V_1UV_2U\{S\}, \Sigma_1 U \Sigma_2, R_1UR_2U\{S\rightarrow S_1S_2\},S)$

(3) The context-free Languages are closed under Kleene star:

• If L_1 is a context free language then there exists a context-free grammar $G_1=(V_1,\Sigma_1,R_1,S_1)$ such that $L_1= L(G_1)$.

• We will build a new grammar G such that $L(G)=L(G_1)^*$ G will contain all the rules of G_1.

• We add to G a new start symbol S and two new rules. $S\rightarrow Ɛ$ and $S\rightarrow SS_1$

So $G = (V_1 U \{S\}, \Sigma_1, R_1U \{S\rightarrow Ɛ, S\rightarrow SS_1), S)$

(4) The context-free languages are closed under reverse

• If L is a context free language then it is generated by some Chomsky Normal Form from grammar G= (V,Σ,R, S).

• Every rule in G is of the form X→BC or X→a, where X, B, and C are elements of (V-Σ) and a ϵ Σ

• So construct, from G, a new grammar G^1 , Such that $L(G^1)= L^R$.

• $G^1= (V_G, \Sigma_G, R', S_G)$, Where R' is constructed as follows:

- For every rule in G of the form X→BC, add to R' the rule X→CB
- For every rule in G of the form X→ a then add to R' the rule X→ a

(5) The context-free languages are closed under letter Substitution

• Consider two alphabets Σ_1 and Σ_2 .

•Let ***sub*** be any function from Σ_1 to Σ_2^*.

•Then ***letsub*** is a letter substitution function from L_1 to L_2 iff **letsub(L_1)** ={ w∈ Σ_2^*: ∃y ∈ L1 (w=y except that every character c of y has replaced by sub(c))}.

Example : Let y= VTU ∈ L_1 And sub(c) is given as : sub(V) = Visvesvaraya

sub(T) = Technological

sub(U)= University

Then , sub(VTU) = Visvesvaraya Technological University

Closure Under Intersection, Complement, and Difference

Theorem:The Context-free language are not closed under intersection, complement or difference.

1) The context-free languages are not closed under intersection

The proof is by counter example. Let: $L_1=\{a^nb^nc^m|n,m\geq 0\}$ $L_2=\{a^mb^nc^n|n,m\geq 0\}$ Both $L_1$and L_2 are context-free since there exist straight forward CFGs for them.

But now consider: $L = L_1 \cap L_2 = \{ a^nb^nc^n \mid n,m\geq 0\}$. If the context-free languages were closure under intersection. L would have to be context-free. But we have proved that L is not CFG by using pumping lemma for CFLs.

(2) The context-free languages are not closure under

Given any sets $L_1$and L_2 , $L_1 \cap L_2 = \neg(\neg L_1 \cup \neg L_2)$

• The context-free languages are closed under union.

• But we just showed that they are not, thus they are not closed under complement either.

• So, if they were also closed under complement, they would necessarily be closed under intersection.

(3) The context-free languages are not closed under difference (subtraction) :

Given any language L and $\neg L=\Sigma^*- L$.

Σ^* is context-free So, if the context-free languages were closed under difference, the complement of any CFL would necessarily be context-free But we just showed that is not so.

Closure Under Intersection With the Regular Languages

Theorem: The context-free languages are closed under intersection with the regular languages.

Proof: The proof is by construction.

• If L_1 is context-free, then there exists some PDA $M_1 = (K_1, \Sigma, \Gamma_1, \Delta_1, S_1, A_1)$ that accepts it.

• If L_2 is regular then there exists a DFSM $M_2 = (K_2, \Sigma, \delta, S_2, A_2)$ that accepts it.

• We construct a new PDA, M_3 that accepts $\mathbf{L_1 \cap L_2}$. M_3 will work by simulating the parallel execution of M_1 and M_2.

• $M_3 = (K_1 X K_2, \Sigma, \Gamma_1, \Delta_3, (S_1, S_2), A_1 X A_2)$, Where Δ_3 is built as follows:

- For each transition $((q_1, a, \beta), (p_1, y))$ in Δ_1 and each transition $((q_2, a), p_2)$ in δ, add Δ_3 the transition: $(((q_1,q_2), a, \beta), ((p_1,p_2), y))$.
- For each transition $((q_1, \varepsilon, \beta), (p_1, y))$ in Δ_1 and each state q_2 in k_2, add to Δ_3 the transition: $(((q_1,q_2), \varepsilon, \beta), ((p_1,p_2), y))$.

Closure Under Difference with the Regular Language.

Theorem: The difference (L_1-L_2) between a context-free language L_1 and a regular language L_2 is context-free.

Proof: $L_1 - L_2 = L_1 \cap \neg L_2$

• If L_2 is regular, then, since the regular languages are closed under complement, $\neg L_2$ is also regular.

• Since L_1 is context-free, by Theorem we already proved that $L_1 \cap \neg L_2$ is context-free.

Using the Pumping Theorem in Conjunction with the Closure Properties

Languages that impose no specific order constraints on the symbols contained in their strings are not always context-free. But it may be hard to prove that one isn't just by using the Pumping Theorem. In such a case it is proved by considering the fact that the context-free languages areclosed under intersection with the regular languages.

Deterministic Context-Free Languages

The technique used to show that the regular languages are closed under complement starts with a given (possibly nondeterministic) FSM M_1, we used the following procedure to construct a newFSM M_2 such that $L(M_2) = \neg L(M_1)$:

The regular languages are closed under complement, intersection and difference. Why are the context-free languages different? Because the machines that accept them may necessarily be nondeterministic.

1. From M_1, construct an equivalent DFSM M', using the algorithm ***ndfsmtodfsm***, presented in the proof of Theorem5.3. (If $M_1$is already deterministic. M'=M_1.)
2. M' must be stated completely. so if it is described with an implied dead state, add the dead state and all required transitions to it.
3. Begin building M_2 by setting it equal to M'. Then swap the accepting and the non-accepting states. So M_2 M'= ($K_{M'}$,.Σ,$\delta_{M'}$,$s_{M'}$,$K_{M'}$-$A_{M'}$).

We have no PDA equivalent of **ndfstodfsm** because there provably isn't one. We defined a PDA M to be deterministic iff:

- Δ_M contains opairs of transitions that compete with each other, and
- if q is an accepting state of M, then there is no transition ((q,ε,ε),(p,a)) for any p or a.

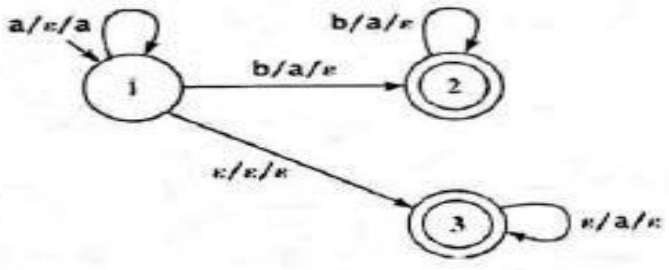
What is a Deterministic Context-Free language?

- Let $ be an end-of-string marker. A language L is deterministic context-free iff L$ can be accepted by some deterministic PDA.

EXAMPLE: Why an End-of-String Marker is Useful

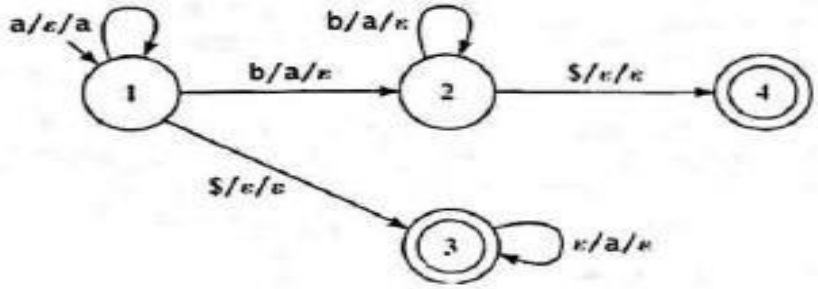
Let L= a* U { $a^n b^n$| n>0}

Consider any PDA M that accepts L. When it begins reading a's, M must push them onto the stack in case there are going to be b's. But if it runs out of input without seeing b's, it needs a way to pop those a's from the stack before it can accept. Without an end-of-string marker, there is no way to allow that popping to happen only when all the input has been read.



For example, the PDA accepts L, but it is nondeterministic because the transition to state3 (where the a's will be popped) can compete with both of the other transitions from state1.

With an end-of-string marker, we can build the deterministic PDA, which can only take the transition to state3, the a-popping state. When it sees the $:



NOTE: Adding the end-of-string marker cannot convert a language that was not context-free into one that is.

CFLs and Deterministic CFLs

Theorem: Every deterministic context-free language is context-free.

Proof:

If L is deterministic context-free, then L$ is accepted by some deterministic PDA M=(K,Σ,*F*,,s,A) . From M, we construct M' such that L (M') = L. We can define the following procedure to construct M': without$(M:PDA)=

1.Initially. set M' to M.

/*Make the copy that does not read any input.

2.For every state q in M, add to M' a new state q'.

3.For every transition $((q, \varepsilon, \gamma_1),(p,\gamma_2))$ in M do:

3.1. Add to $_M$ the transition $((q',\varepsilon,\gamma_1),(p',\gamma_2))$.

/*Link up the two copies.

4.For every transition $((q,\$,\gamma_1),(p,\gamma_2))$ in M do:

4.1. Add to $_{M'}$ the transition $((q,\varepsilon,\gamma_1),(p',\gamma_2))$.

4.2. Remove $((q,\$,\gamma_1),(p,\gamma_2))$ from $_{M'}$

/*Set the accepting state s of M'.

5.$A_{M'} = \{q':q \in A\}$.

Closure Properties of the Deterministic Context-Free Languages

1) Closure Under Complement

Theorem: The deterministic context-free languages are closed under complement.

Proof: The proof is by construction. If L is a deterministic context-free language over the alphabet Σ, then L$ is accepted by some deterministic PDA M = (K, Σ U{$}, *F*, , s, A).

We need to describe an algorithm that constructs a new deterministic PDA that accepts (¬L)$.

We defined a construction that proceeded in two steps:

- Given an arbitrary FSM, convert it to an equivalent DFSM, and then
- Swap accepting and non accepting states.

A deterministic PDA may fail to accept an input string *w* for any one of several reasons:

1. Its computation ends before it finishes reading *w*.
2. Its computation ends in an accepting state but the stack is not empty.
3. Its computation loops forever, following ε-transitions, without ever halting in either an accepting or a non accepting state.
4. Its computation ends in a non accepting state.

If we simply swap accepting and non accepting states we will correctly fail to accept every string that M would have accepted (i.e., every string in L$). But we will not necessarily accept every string in (¬L)$. To do that, we must also address issues 1through 3 above.

An additional problem is that we don't want to accept ¬L(M). That includes strings that do not end in $. We must accept only strings that do end in $ and that are in (¬L)$.

2) Non closure Under Union

Theorem: The deterministic context-free languages are not closed under union.

Proof: We show a counter example:

Let, L_1={ $a^i b^j c^k$| i,j,k ≥0 and i≠ j } and L_2={ $a^i b^j c^k$ | i,j,k≥0 and j ≠ k}

Let, L' = L_1U L_2= { $a^i b^j c^k$ | i,j,k≥0 and ((i≠ j) and (j ≠ k)) }.

Let, L'' = ¬L'.

$=\{ a^i b^j c^k \mid i,j,k \geq 0 \text{ and } (i=j=k)\} \cup \{w \in \{a,b,c\}^* : \text{the letters are out of order}\}$.

Let, $L''' = L'' \cap a^*b^*c^* = \{a^n b^n c^n \mid n \geq 0\}$

But L''' is $A^n B^n C^n = \{ a^n b^n c^n \mid n \geq 0\}$, which we have shown is not context-free.

3) Non Closure Under Intersection

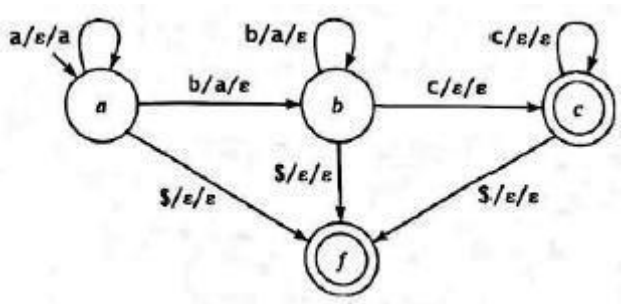
Theorem: The deterministic context-free languages are not closed under inter section.

Proof: We show a counter example:

Let, $L_1 = \{ a^i b^j c^k \mid i,j,k \geq 0 \text{ and } i = j \}$ and $L_2 = \{ a^i b^j c^k \mid i,j,k \geq 0 \text{ and } j = k\}$

Let, $L' = L_1 \cap L_2 = \{ a^n b^n c^n \mid n \geq 0\}$

L_1 and L_2 are deterministic context-free. The deterministic PDA shown accepts $L_1\$$, A similar one accepts L_2. But we have shown that their intersection L' is not context-free much less deterministic context-free.



A hierarchy within the class of context-free languages

Some CFLs are not Deterministic

Theorem: The class of deterministic context-free languages is a proper subset of the class of context-free languages. Thus there exist nondeterministic PDAs for which no equivalent deterministic PDA exists.

Proof: We show that there exists at least one context-free language that is not deterministic context-free.

Consider $L = \{ a^i b^j c^k \mid i,j,k \geq 0 \text{ and } ((i \neq j) \text{ or } (j \neq k)) \}$. L is context-free.

If L were deterministic context-free, then, its complement

$L' = \{ a^i b^j c^k \mid i,j,k \geq 0 \text{ and } (i=j=k) \} \cup \{w \in \{a,b,c\}^* : \text{the letters are out of order}\}$

Would also be deterministic context-free and thus context-free. If L' were context-free, then $L'' = L' \cap a^*b^*c^*$ would also be context-free (since the context-free languages are closed under inter section with the regular languages).

But $L'' = A^n B^n C^n = \{a^n b^n c^n \mid n \geq 0\}$, which is not context free. So

L is context-free but not deterministic context-free.

Since L is context-free, it is accepted by some (non deterministic) PDA M. M is an example of an on deterministic PDA for which no equivalent deterministic PDA L exists. If such a deterministic PDA did exist and accept L, it could be converted into a deterministic PDA that accepted L$. But, if that machine existed. L would be deterministic context-free and we just showed that it is not.

Inherent Ambiguity versus Non determinism

There are context-free languages for which unambiguous grammars exist and there are others that are inherently ambiguous, by which we mean that every corresponding grammar is ambiguous.

Example:

The language $L_1 = \{a^i b^j c^k \mid i, j, k \geq 0 \text{ and } ((i = j) \text{ or } (j = k))\}$ can also be described as $\{a^n b^n c^m | n,m \geq 0\} \cup \{a^n b^m c^m | n,m \geq 0\}$. L_1 is inherently ambiguous because every string that is also in $A^n B^n C^n = \{a^n b^n c^n | n \geq 0\}$ is an element of both sub languages and so has at least two derivations in any grammar for L_1.

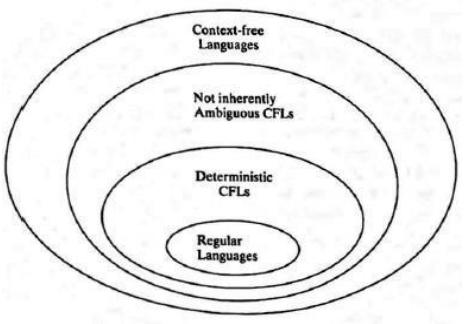
• Now consider the language $L_2 = \{a^n b^n c^m d | n,m \geq 0\} \cup \{a^n b^m c^m e | n,m \geq 0\}$ is not inherently ambiguous.

• Any string in is an element of only one of them (since each such string must end in d or e but not both).

There exists no PDA that can decide which of the two sublanguages a particular string is in until it has consumed the entire string.

What is the relationship between the deterministic context-free languages and the languages that are not inherently ambiguous?

The answer is shown in below Figure.



There exist deterministic context-free languages that are not regular. One example is $A^n B^n = \{a^n b^n | n,m \geq 0\}$.

•There exist context-free languages and not inherently ambiguous. Examples:

(a) $PalEven = \{ww^R : w \in \{a,b\}^*\}$.

(b) $\{a^n b^n c^m d | n,m \geq 0\} \cup \{a^n b^m c^m e | n,m \geq 0\}$.

•There exist languages that are in the outer donut because they are inherently ambiguous. Two examples are:

- $\{a^i b^j c^k | i,j,k \geq 0 \text{ and } ((i \neq j) \text{or} (j \neq k))\}$
- $\{a^i b^j c^k | i,j,k \geq 0 \text{ and } ((i = j) \text{ or } (j = k))\}$

Regular Language is Deterministic Context-Free

Theorem: Every regular language is deterministic context-free.

Proof: The proof is by construction. {$} is regular. So, if L is regular then so is L$ (since the regular languages are closed under concatenation).So there is a DFSM M that accepts it. Using the construction to show that every regular language is context-free Construct, from M a PDA P that accepts L$. P will be deterministic.

Every Deterministic CFL has an Unambiguous Grammar

Theorem: For every deterministic context-free language there exists an unambiguous grammar.

Proof: If a language L is deterministic context-free, then there exists a deterministic PDA M that accepts L$. We prove the theorem by construction of an unambiguous grammar G such that L(M)= L(G). We construct G as follows:

The algorithm PDAtoCFG proceeded in two steps:

1. Invoke ***convenPDAtorestricted(M)*** to build M', an equivalent PDA in restricted normal form.
2. Invoke ***buildgrammar(M'),*** to build an equivalent grammar G

So the construction that proves the theorem is:

buildunambiggrammar(M:deterministicPDA)=

1. Let G=*buildgrammar*(***convertPDAtodetnormalform*****(M)**).
2. Let G' be the result of substituting ε for $ in each rule in which $ occurs.
3. Return G'.

NOTE: The algorithm ***convertPDAtodetnormalform***, is described in the theorem that proves the deterministic context-free languages are closed under complement.

The Decidable Questions

Membership

"Given a language L and a string *w*, is *w* in L?'

This question can be answered for every context-free language and for every context-free language L there exists a PDA M such that M accepts L. But existence of a PDA that accepts L does not guarantee the existence of a procedure that decides it.

It turns out that there are two alternative approaches to solving this problem, both of which work:

- **Use a grammar:** Using facts about every derivation that is produced by a grammar in Chomsky normal form, we can construct an algorithm that explores a finite number of derivation paths and finds one that derives a particular string w iff such a path exists.
- **Use a PDA** : While not all PDAs halt, it is possible, for any context-free language L, to craft a PDA M that is guaranteed to halt on all inputs and that accepts all strings in L and rejects all strings that are not in L.

Using a Grammar to Decide

Algorithm for deciding whether a string *w* is in a language L:

***decideCFLusingGrammar*(L: CFL,*w*: string) =**

1. If L is specified as a PDA, use ***PDA to CFG***, to construct a grammar G such that L(G) =L (M).

2. If L is specified as a grammar G, simply use G.

3. If w =ε then if S_G is nullable then accept, otherwise reject.

4. If w ≠ ε then:

   4.1. From G, construct G' such that L (G') = L(G)-{ε} and G' is in Chomsky normal form.

   4.2. If G derives to, it does so in (2 • |*w*| - 1) steps. Try all derivations in G of that number of steps. If one of them derives w, accept. Otherwise reject.

Using a PDA to Decide

A two-step approach.

- We first show that, for every context-free language L, it is possible to build a PDA that accepts L-{ε} and that has no ε-transitions.
- Then we show that every PDA with no ε-transitions is guaranteed to halt

Elimination of ε-Transitions

Theorem: Given any context-free grammar G=(V,Σ,R,S), there exists a PDA M such that L(M)=L(G)-{ε} and M contains no transitions of the form

((q1,ε,α),(q2,β)). In other words, every transition reads exactly one input character.

Proof: The proof is by a construction that begins by converting G to Greibach normal form. Now consider again the algorithm ***cfgtoPDAtopdown***, which builds, from any context-free grammar G, a PDA M that, on input w, simulates G deriving w, starting from S.

M= ({p,q},Σ,V,Δ, p,{q}), where Δ contains:

1. The start-up transition ((p,ε,ε),(q,S)), which pushes the start symbol on to the stack and goes to state q.
2. For each rule $X \rightarrow s_1 s_2 ... s_n$, in R, the transition ((q,ε,X),(q,$s_1 s_2 ... s_n$)), which replaces X by $s_1 s_2 ... s_n$. If n=0 (i.e., the right-hand side of the rule is ε), then the transition ((q, ε, X), (q, ε)).
3. For each character c ∈ Σ. the transition ((q, c, c), (q,ε)), which compares an expected character from the stack against the next input character.

If G contains the rule $X \rightarrow c s_2 ... s_n$, (where *c* ∈Σ and s_2 through s_n, are elements of V-Σ), it is not necessary to push *c* onto the stack, only to pop it with a rule from step 3.

Instead, we collapse the push and the pop into a single transition. So we create a transition that can be taken only if the next input character is *c*. In that case, the string s_2 through s_n is pushed onto the stack.

Since terminal symbols are no longer pushed onto the stack. We no longer need the transitions created in step3 of the original algorithm.

So, M=({p,q},Σ,V,Δ,p,{q}), where Δ contains:

1. The start-up transitions: For each rule $S \rightarrow c s_2 ... s_n$ the transition ((p,c,ε),(q,$s_2 ... s_n$)).
2. For each rule $X \rightarrow c s_2 ... s_n$ (where c∈Σ and s_2 through s_n, are elements of V-Σ), the transition ((q,c,X),(q,$s_2 ... s_n$)).

cfgtoPDAnoeps(G:context-freegrammar)=

1. Convert G to Greibach normal form, producing G'.
2. From G' build the PDA M described above.

Halting Behavior of PDAs Without ε-Transitions

Theorem: Let M be a PDA that contains no transitions of the form ((q_1,ε,s1),(q_2,s_2)). i.e., no ε-transitions. Consider the operation of M on input w∈Σ*. M must halt and either accept or reject w. Let n=|w|.

We make three additional claims:

a) Each individual computation of M must halt within *n* steps.

b) The total number of computations pursued by M must be less than or equal to b^n, where *b* is the maximum number of competing transitions from any state in M.

c) The total number of steps that will be executed by all computations of M is bounded by nb^n

Proof:

a) Since each computation of M must consume one character of w at each step and M will halt when it runs out of input, each computation must halt within n steps.

b) M may split into at most b branches at each step in a computation. The number of steps in a computation is less than or equal to n. So the total number of computations must be less than or equal to b^n.

c) Since the maximum number of computations is b^n and the maximum length of each is n, the maximum number of steps that can be executed before all computations of M halt is nb^n.

So a second way to answer the question, "Given a context-free language L and a string w, is w in L?" is to execute the following algorithm:

decideCFLusingPDA(L:CFL,w:string)=

1. If L is specified as a PDA, use ***PDAtoCFG,*** to construct a grammar G such that L(G)=L(M).
2. If L is specified as a grammar G, simply use G.
3. If w=ε then if S_G is nullable then accept, otherwise reject.
4. If w≠ε then:
    4.1. From G, construct G' such that L(G')=L(G)-{ε} and G' is in Greibach normal form.
    4.2. From G' construct, using ***cfgtoPDAnoeps***, a PDA M' such that L(M')=L(G') and M' has no ε-transitions.
    4.3. We have proved previously that, all paths of M' are guaranteed to halt within a finite number of steps. So run M' on w, Accept if M' accepts and reject otherwise.

Emptiness and Finiteness

Decidability of Emptiness and Finiteness

Theorem: Given a context-free language L. There exists a decision procedure that answers each of the following questions:

1. Given a context-free language L, is L=⦰?
2. Given a context-free language L, is L infinite?

Since we have proven that there exists a grammar that generates L iff there exists a PDA that accepts it. These questions will have the same answers whether we ask them about grammars or about PDAs.

Proof :

***decideCFLempty*(G: context-free grammar) =**

1. Let G' =***removeunproductive(G).***

2. If S is not present in G' then return True else return False.

***decideCFLinfinite*(G:context-freegrammar)=**

1. Lexicographically enumerate all strings in Σ* of length greater than b^n and less than or equal to $b^{n+1}+b^n$.
2. If, for any such string w, ***decideCFL*(L,w)** returns *True* then return *True*. L is infinite.
3. If, for all such strings w, ***decideCFL*(L,w)** returns *False* then return *False*. L is not infinite.

The Undecidable Questions

- Given a context-free language L, is L=Σ*?
- Given a CFL L, is the complement of L context-free?
- Given a context-free language L, is L regular?
- Given two context-free languages L_1 and L_2 is $L_1=L_2$?
- Given two context-free languages L_1 and L_2, is $L_1 \subseteq L_2$?
- Given two context-free languages $L_1$and L_2, is $L_1 \cap L_2 = \varnothing$?
- Given a context-free language L, is L inherently ambiguous?
- Given a context-free grammar G, is G ambiguous?

TURING MACHINE

The Turing machine provides an ideal theoretical model of a computer. Turing machines are useful in several ways:

- Turing machines are also used for determining the undecidability of certain languages and
- As an automaton, the Turing machine is the most general model, It accepts type-0 languages.
- It can also be used for computing functions. It turns out to be a mathematical model of partial recursive functions.
- Measuring the space and time complexity of problems.

Turing assumed that while computing, a person writes symbols on a one-dimensional paper (instead of a two dimensional paper as is usually done) which can be viewed as a tape divided into cells. In

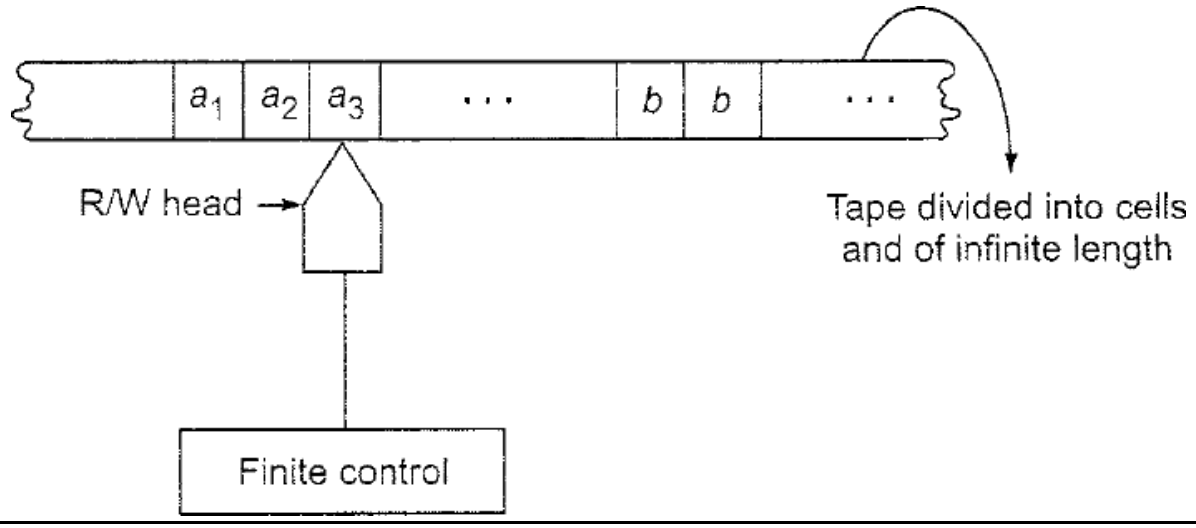
Turing machine one scans the cells one at a time and usually performs one of the three simple operations, namely:

(i) Writing a new symbol in the cell being currently scanned,

(ii) Moving to the cell left of the present cell, and

(iii) Moving to the cell right of the present cell.

Turing machine model



•Each cel can store only one symbol.

•The input to and the output from the finite state automaton are affected by the R/W head which can examine one cell at a time.

In one move, the machine examines the present symbol under the R/W head on the tape and the present state of an automaton to determine:

(i) A new symbol to be written on the tape in the cell under the R/W head,

(ii) A motion of the R/W head along the tape: either the head moves one cell left (L),or one cell right (R).

(iii) The next state of the automaton, and

(iv) Whether to halt or not.

Definition:

Turing machine M is a 7-tuple, namely $(Q, \Sigma, \Gamma, , q_0, b, F)$, where

1. Q is a finite nonempty set of states.
2. Γ is a finite nonempty set of tape symbols,
3. $b \in \Gamma$ is the blank.
4. Σ is a nonempty set of input symbols and is a subset o f Γ and $b \notin \Sigma$.

5. is the transition function mapping (q,x) onto (q',y,D) where D denotes the direction of movement of R/W head; D=L orR according as the movement is to the left or right.
6. $q_0 \in Q$ is the initial state, and
7. $F \subseteq Q$ is the set of final states.

Notes:

(1) The acceptability of a string is decided by the reachability from the initial state to some final state.

(2) may not be defined for some elements of $Q \times \Gamma$.

REPRESENTATION OF TURINGMACHINES

We can describe a Turing machine employing

(i) Instantaneous descriptions using move-relations.

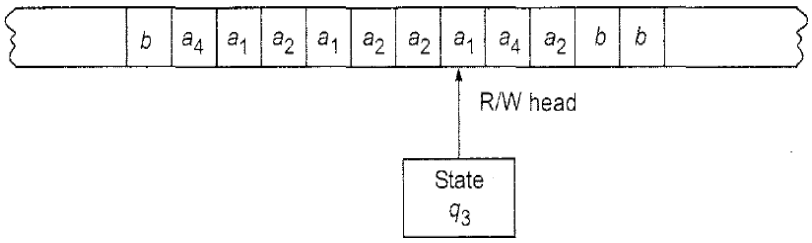
(ii) Transition table, and

(iii) Transition diagram (Transition graph).

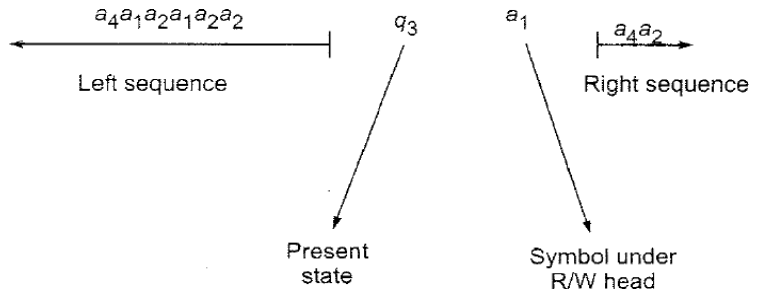
REPRESENTATION BY INSTANTANEOUS DESCRIPTIONS

Definition: An ID of a Turing machine M is a string $\alpha\beta\gamma$, where β is the present state of M, the entire input string is split as $\alpha\gamma$, the first symbol of γ is the current symbol ***a*** under the R/W head and γ has all the subsequent symbols of the input string, and the string α is the substring of the input string formed by all the symbols to the left of ***a***.

EXAMPLE: A snapshot of Turing machine is shown in below Fig. Obtain the instantaneous description.



The present symbol under the R/W head is a_1. The present state is q_3. So a_1 is written to the right of q_3 The nonblank symbols to the left of al form the string $a_4a_1a_2a_1a_2a_2$, which is written to the left of q_3. The sequence of nonblank symbols to the right of a_1 is a_4a_2. Thus the ID is as given in below Fig.



Notes: (1) For constructing the ID, we simply insert the current state in the input string to the left of the symbol under the R/W head.

(2) We observe that the blank symbol may occur as part of the left or right substring.

REPRESENTATION BY TRANSITION TABLE

We give the definition of in the form of a table called the transition table If (q, a)=(γ,α,β). We write $\alpha\beta\gamma$ under the α-column and in the q-row. So if we get $\alpha\beta\gamma$ in the table, it means that α is written in the current cell, β gives the movement of the head (L or R) and γ denotes the new state into which the Turing machine enters.

EXAMPLE:

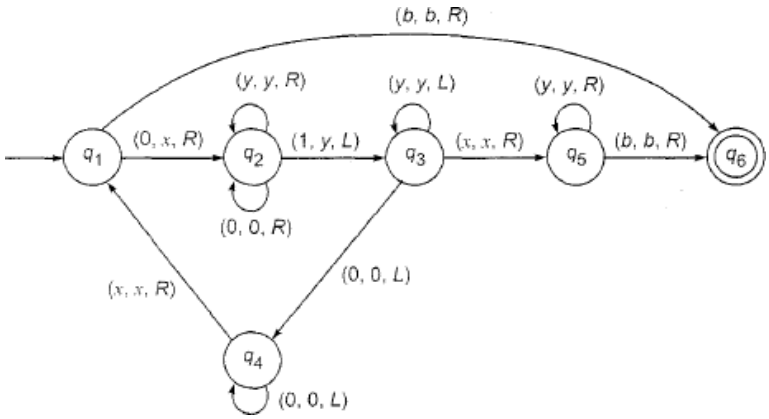
Consider, for example, a Turing machine with five states $q_1,...,q_5$ where q_1 is the initial state and q_5 is the (only) final state. The tape symbols are 0,1and b. The transition table given below describes :

Present state	Tape symbol		
	b	0	1
$\rightarrow q_1$	$1Lq_2$	$0Rq_1$	
q_2	bRq_3	$0Lq_2$	$1Lq_2$
q_3		bRq_4	bRq_5
q_4	$0Rq_5$	$0Rq_4$	$1Rq_4$
(q_5)	$0Lq_2$		

REPRESENTATION BY TRANSITION DIAGRAM (TD)

The states are represented by vertices. Directed edges are used to represent transition of states. The labels are triples of the form (α,β,γ)where $\alpha,\beta \in \Gamma$and$\gamma \in$ {L,R}.When there is a directed edge from q_i to q_j with label (α,β,γ),it means that (q_i,α)=(q_j,β,γ).

EXAMPLE:



LANGUAGE ACCEPTABILITY BY TURING MACHINES

Let us consider the Turing machine M=(Q,Σ,**Γ**,,q0,b,F). A string *w* in Σ* is said to be accepted by M, if $q_0w \overset{*}{\vdash} \alpha_1 p \alpha_2$ for some P∈F and $\alpha_1, \alpha_2 \in \Gamma^*$.

EXAMPLE: Consider the Turing machine M described by the table below

Present state	Tape symbol				
	0	1	x	y	b
→ q_1	xRq_2				bRq_5
q_2	$0Rq_2$	yLq_3		yRq_2	
q_3	$0Lq_4$		xRq_5	yLq_3	
q_4	$0Lq_4$		xRq_1		
q_5				$yxRq_5$	bRq_5
(q6)					

IDs for the strings (a) 011 (b)0011 (c)001

(a) $q_1011 \vdash xq_211 \vdash q_3xy1 \vdash xq_5y1 \vdash xyq_51$

As (q_5,1) is not defined, M halts; so the input string 011 is not accepted

(b) $q_10011 \vdash xq_2011 \vdash x0q_211 \vdash xq_30y1 \vdash q_4x0y1 \vdash xq_10y1 \vdash xxq_2y1 \vdash xxyq_21 \vdash xxq_3yy \vdash xq_3xyy \vdash xxq_5yy \vdash xxyq_5y \vdash xxyyq_5b \vdash xxyybq_6$

M halts. As q_6 is an accepting state, the input string 0011is accepted byM.

(c) $q_1001 \vdash xq_201 \vdash x0q_21 \vdash xq_30y \vdash q_4x0y \vdash xq_10y \vdash xxq_2y \vdash xxyq_2$

M halts. As q_2 is not an accepting state,001 is not accepted by M.

DESIGN OF TURING MACHINES

Basic guidelines for designing a Turing machine:

1. **The fundamental objective in scanning a symbol by the R/W head is to know what to do in the future.** The machine must remember the past symbols scanned. The Turing machine can remember this by going to the next unique state.
2. **The number of states must be minimized**. This can be achieved by changing the states only when there is a change in the written symbol or when there is a change in the movement of the R/W head.

EXAMPLE 1

Design aTuring machine to recognize all strings consisting of an even number of 1's.

Solution:

The construction is made by defining moves in the following manner:

(a) q_l is the initial state. M enters the state q_2 on scanning 1 and writes b.

(b) If M is in state q_2 and scans 1, it enters q_l and writes b.

(c) q_l is the only accepting state.

Symbolically M= ($\{q_l,q_2\},\{1,b\},\{1,b\},,q,b,\{q_l\}$), Where is defined by ,

Present state	1
→ q_1	bq_2R
q_2	bq_1R

Let us obtain the computation sequence of 11:

$q_111 \vdash bq_21 \vdash bbq_1$

As q_1 is an accepting state 11 is accepted.

Let us obtain the computation sequence of 111:

$q_1111 \vdash bq_211 \vdash bbq_11 \vdash bbbq_2$

As q_2 is an not accepting state 111 is not accepted.

EXAMPLE 2: Design a TM that accepts $\{0^n1^n | n \geq 0\}$

Solution: We require the following moves:

(a) If the leftmost symbol in the given input string *w* is 0, replace it by *x* and move right till we encounter a leftmost 1in *w*. Change it to *y and* move backwards.

(b) Repeat (a) with the leftmost 0. If we move back and forth and no 0 or 1 remains. Move to a final state.

(c) For strings not in the form 0^n1^n, the resulting state has to be non-final.

we construct a **TM *M as follows:*** $M = (Q, \Sigma, \Gamma, , q_0, b, F)$

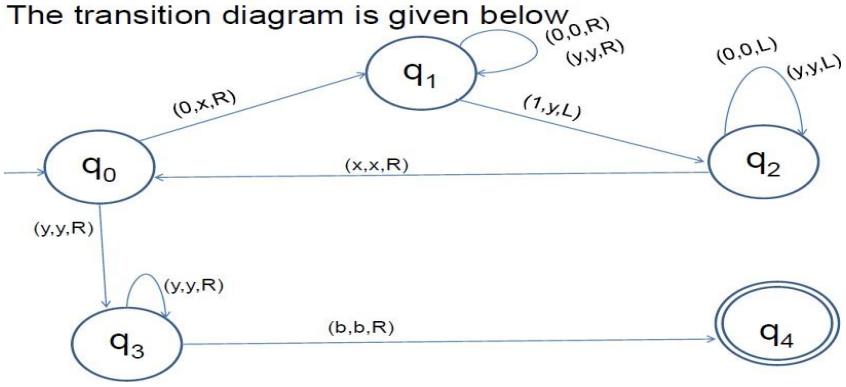
$Q = \{q_0, q_1, q_2, q_3, q_f\}$

$F = \{q_f\}$

$\Sigma = \{0,1\}$

$\Gamma = \{0,1,x,y,b\}$

The transition diagram is given below



Computation sequence of 0011:

$$q_0 0011 \vdash xq_1 011 \vdash x0q_1 11 \vdash xq_2 0y1 \vdash q_2 x0y1 \vdash xq_0 0y1$$
$$\vdash xxq_1 y1 \vdash xxyq_1 1 \vdash xxq_2 yy \vdash xq_2 xyy \vdash xxq_0 yy \vdash xxyq_3 y$$
$$\vdash xxyyq_3 = xxyyq_3 b \vdash xxyybq_4 b$$

q_4 is final state, hence 0011 is accepted by M.

TECHNIQUES FOR TM CONSTRUCTION

1. TURING MACHINE WITH STATIONARY HEAD

Suppose, we want to include the option that the head can continue to be in the same cell for some input symbol. Then we define (q,a) as (q',y,S).This means that the TM, on reading the input symbol a, changes the state to q' and writes y in the current cell in place of a and continues to remain in the same cell. In this model (q, a) =(q', y, D) where D = L, R or S.

2. STORAGE IN THE STATE

We can use a state to store a symbol as well. So the state becomes a pair*(q,a)* where q is the state and a is the tape symbol stored in (q, a). So the new set of states becomes Qx**Γ**.

EXAMPLE: Construct a TM that accepts the language 0 1* + 1 0*.

We have to construct a TM that remembers the first symbol and checks that it does not appear afterwards in the input string.

So we require two states, q_0, q_1. The tape symbols are 0,1 and b. So the TM, having the '**storage facility in state**', is M=({q_0,q_1}X{0,1,b},{0,1},{0,1,b},,[q_0,b],[q_1,b]})

We describe by its implementation description.

1. In the initial state, M is in q_0 and has b in its data portion. On seeing the first symbol of the input sting w, M moves right, enters the state q_1 and the first symbol, say a, it has seen.

2. M is now in [q_1,a].

(i) If its next symbol is b, M enters [q_1,b], an accepting state.

(ii) If the next symbol is a, M halts without reaching the final state (i.e. is not defined).

(iii) If the next symbol is ā, (ā=0 if a=1 and ā=1 if a=0), M moves right without changing state.

3. Step2 is repeated until M reaches [q_1,b] or halts (is not defined for an input symbol in w).

3. MULTIPLE TRACK TURING MACHINE

In a multiple track TM, a single tape is assumed to be divided into several tracks. Now the tape alphabet is required to consist of k-tuples of tape symbols, k being the number of tracks. In the case of the standard Turing machine, tape symbols are elements of r; in the case of TM with multiple tracks, it is $\mathbf{\Gamma}^k$.

4. SUBROUTINES

First a TM program for the subroutine is written. This will have an initial state and a 'return' state. After reaching the return state, there is a temporary halt for using a subroutine, new states are introduced. When there is a need for calling the subroutine, moves are effected to enter the initial state for the subroutine. When the return state of the subroutine is reached, return to the main program of TM.

EXAMPLE: Design a TM which can multiply two positive integers.

Solution: The input (m,n), m,n being given ,the positive integers represented by 0^m10^n. M starts with 0^m10^n in its tape. At the end of the computation, 0^{mn} (mn in unary representation) surrounded byb's is obtained as the output.

The major steps in the construction are as follows:

1. 0^m10^n1 is placed on the tape (the output will be written after the rightmost 1).
2. The leftmost 0 is erased.
3. A block of n 0's is copied onto the right end.
4. Steps 2 and 3 are repeated m times and 10^m10^{mn} is obtained on the tape.
5. The prefix 10^m1 of 10^m10^{mn} is erased, leaving the product 0^{mn} as the output.

For every 0 in 0^m, 0^n is added onto the right end. This requires repetition of step3. We define a subroutine called COPY for step3. For the subroutine COPY the initial state is q_1 *and the final state is* q_5 is given by the transition table as below:

The transition table for the **SUBROUTINE COPY**

State	Tape symbol			
	0	1	2	b
q_1	$q_2 2R$	$q_4 1L$	—	—
q_2	$q_2 0R$	$q_2 1R$	—	$q_3 0L$
q_3	$q_3 0L$	$q_3 1L$	$q_1 2R$	—
q_4	—	$q_5 1R$	$q_4 0L$	—
q_5	—	—	—	—

The Turing machine M has the initial state q_0. The initial ID for M is $q_00^m10^n$. O^n seeing 0,the following moves take place

$$q_00^m10^n1 \vdash bq_60^{m-1}10^n1 \vdash^* b0^{m-1}q_610^n1 \vdash b0^{m-1}1q_10^n1$$

q_1 is the initial state of COPY. The following moves take place for M_1:

$$q_10^n1 \vdash 2q_20^{n-1}1 \vdash^* 20^{n-1}1q_3b \vdash 20^{n-1}q_310 \vdash^* 2q_10^{n-1}10$$

After exhausting 0s, q_1 encounters 1. M_1 moves to state q_4. All 2's are converted back to 0's and M_1 halts in q_5. The TM M picks up the computation by starting from q_5 The q_0 and q_6 are the states of M. Additional states are created to check whether reach 0 in 0^m gives rise to 0^m at the end of the rightmost 1 in the input string. Once this is over, M erases 10^n1 and finds 0^{mn} in the input tape.

M can be defined by M=($\{q_0,q_1,....q_{12}\}\{0,1\},\{0,,2,b\},,q_0,b,\{q_{12}\}$) where is defined by table givenbelow:

	0	1	2	b
q_0	q_6bR	—	—	—
q_6	q_60R	q_11R	—	—
q_5	q_70L	—	—	—
q_7	—	q_81L	—	—
q_8	q_90L	—	—	$q_{10}bR$
q_9	q_90L	—	—	q_0bR
q_{10}	—	$q_{11}bR$	—	—
q_{11}	$q_{11}bR$	$q_{12}bR$	—	—

ADDITIONAL PROBLEMS

1. Design a Turing machine to obtain complement of a binary number.

IDEA OF CONSTRUCTION:

1) If symbol is 0 change it to 1, move read write head to RIGHT
2) If symbol is 1 change it to 0, move read write head to RIGHT
3) Symbol is b (blank) don't change, move read write head to RIGHT, and HALT.

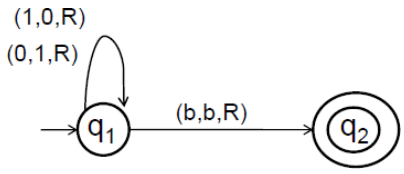
The construction is made by defining moves in the following manner:

(a) q_l is the initial state. On scanning 1, no change in state and write 0 and move head to RIGHT.

(c) If M is in state $q_1$and scans blank, it enters q_2 and writes b move to right.

(d) q_2 is the only accepting state.

Symbolically, M=({q_l,q_2},{1,0,b},{1,0,b},,q_l,b,{q_2}) Where is defined by:



The computation sequence of 1010:

$q_1 1010 \vdash 0q_1 010 \vdash 01q_1 10 \vdash 010q_1 0 \vdash 0101q_1 b$

$\vdash 0101bq_2 b$

2. Design a TM that converts binary number into its 2's complement representation.

IDEA OF CONSTRUCTION:

- Read input from left to right until right end blank is scanned.
- Begin scan from right to left keep symbols as it is until 1 found on input file.

- If 1 found on input file, move head to left one cell without changing input.
- Now until left end blank is scanned, change a**l** 1's to 0 and 0's to 1.

We require the following moves:

(a) Let q_1 be initial state, until blank is scanned, move head to RIGHT without changing anything. On scanning blank, move head to RIGHT change state to q_2 without changing the content of input.

(b) If q_2 is the state, until 1 is scanned, move head to LEFT without changing anything. On reading 1, change state to q_3, move head to LEFT without changing input.

(c) If q_3 is the state, until blank is scanned, move head to LEFT, if symbol is 0 change to 1, otherwise if symbol is 1 change to 0.On finding blank change state to q_4, move head to LEFT without Changing input.

(d) q_4 is the only accepting state.

We construct a **TM *M as follows:***

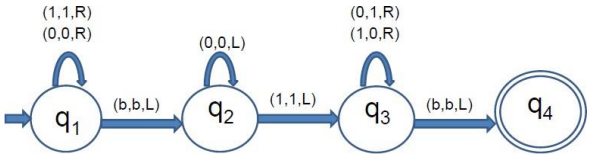
M = (Q, Σ,, , q_0,b, F)

Q = {q_1,q_2,q_3,q_4}

F = {q_4 }

Σ = {0,1}

Γ= {0,1,b}



3.Design a TM that add two integers

IDEA OF CONSTRUCTION:

- Read input from LEFT to RIGHT until blank (separator of two numbers) is found.
- Continue LEFT to RIGHT until blank (end of second number) is found.
- Change separator b to 1 move head to RIGHT.
- move header to Left (to point rightmost 1)
- Change 1 to b and move right, Halt.

We require the following moves:

(a) In q_1 TM skips1's until it reads b (separator),changes to1and goes to q_1

(b) In q_2 TM skips1's until it reads b (end of input), turns left and goes to q_3

(c) In q_3, TM reads 1 and changes to b go to q_4.

(d) q_4 is the final state, TM halts.

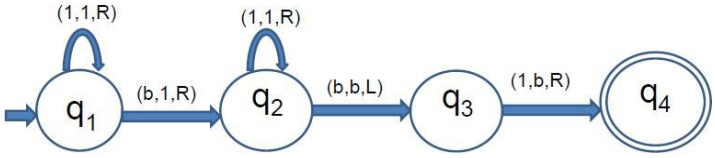
we construct a **TM M as follows:** M = (Q, Σ, **Γ**, , q_0,b, F)

$Q = \{q_1,q_2,q_3,q_4\}$

$F = \{q_4\}$

$\Sigma = \{ b,1\}$

Γ= {1,b}



4. Design a TM that accepts the set of all palindromes over {0,1}*

IDEA OF CONSTRUCTION:

- If it is 0 and changes to X, similarly if it is 1, it is changed to Y, and moves right until it finds blank.
- Starting at the left end it checks the first symbol of the input,
- Nowmovesonestepleftandcheckwhetherthesymbolreadmatchesthemostrecentlychanged.Ifsoiti salsochangedcorrespondingly.
- Now machine moves back left until it finds 0 or 1.
- This process is continued by moving left and right alternately until a**l** 0's and 1's have been matched.

We require the following moves:

1.If state is q0and it scans 0.

- Then go to state q1 and change the 0 to an X,
- move RIGHT over a**l** 0's and 1's, until it finds either X or Y or B
- Now move one step left and change state to q_3
- It verifies that the symbol read is 0, and changes the 0 to X and goes to state q_5.

2. If state is q0 and it scans 1

- Then go to state q_2 and change the 1 to an Y,
- Move RIGHT over a**l** 0's and 1's, until it finds either X or Y or B
  - Now move one step left and change state to q_4
  - It verifies that the symbol read is 1, and changes the 1 to Y and goes to state q_5.

3. If state is q_5

- Move LEFT over a**l** 0's and 1's, until it finds either X or Y.
- Now move one step RIGHT and change state to q_0.

- Now at q_0 there are two cases:

  1. If 0's and 1's are found on input , it repeats the matching cycle just described.
  2. If X's and Y's are found on input, then it changes all the 0's to X and all the 1's to Y's.

     The input was a palindrome of even length, Thus, state changed to q_6.

4.If state is q_3 or q_4

If X's and Y's are found on input, it concludes that: **The input was a palindrome of odd length, t**hus, state changed to q_6.

We construct a **TM *M as follows:***

$M = (Q, \Sigma, \Gamma, , q_0, b, F)$

$Q = \{q_0, q_1, q_2, q_3, q_4, q_5, q_6\}$

$F = \{q_6\}$

$\Sigma = \{ b, 1, 0\}$

$\Gamma = \{X, Y, b\}$

state	0	1	X	Y	B
q_0	(q_1, X, R)	(q_2, Y, R)	(q_6, X, R)	(q_6, Y, R)	(q_6, B, R)
q_1	$(q_1, 0, R)$	$(q_1, 1, R)$	(q_3, X, L)	(q_3, Y, L)	(q_3, B, L)
q_2	$(q_2, 0, R)$	$(q_2, 1, R)$	(q_4, X, L)	(q_4, Y, L)	(q_4, B, L)
q_3	(q_5, X, L)	-	(q_6, X, R)	(q_6, Y, R)	-
q_4	-	(q_5, Y, L)	(q_6, X, R)	(q_6, Y, R)	-
q_5	$(q_5, 0, L)$	$(q_5, 1, L)$	(q_0, X, R)	(q_0, Y, R)	-

PRACTICE PROBLEMS

1. Design a Turing machine to replace a**l** a's with X and a**l** b's with Y.
2. Design a Turing machine to accept $a^n b^m$ n>m.
3. Design a Turing machine to accept $a^n b^n$ n<m.
4. Design a Turing machine to accept (0+1)*00(0+1)* .
5. Design a Turing machine to increment a given input.
6. Design a Turing machine to decrement a given input.
7. Design a Turing machine to subtract two unary numbers.
8. Design a Turing machine to multiply two unary numbers.
9. Design a Turing machine to accept a string 0's followed by a 1.

10. Design a Turing machine to verify if the given binary number is an even number or not.
11. Design a Turing machine to shift the given input by one cell to left.
12. Design a Turing machine to shift the given input to the right by one cell .
13. Design a Turing machine to rotate a given input by one cell.
14. Design a Turing machine to erase the tape.
15. Design a Turing machine to accept $a^n b^n c^n$.
16. Design a Turing machine to accept any string of a's & b's with equal number of a's & b's.
17. Design a Turing machine to accept $a^n b^{2n}$.
18. Design a Turing machine to accept $a^n b^k c^m$: where n=m+k.
19. Design a Turing machine to accept $a^n b^k c^m$: where m=n+k.
20. Design a Turing machine to accept $a^n b^k c^m$: where k=m+n.

Church-Turing thesis-1936

- Any algorithmic procedure that can be carried out by a human or a computer, can also be carried out by a Turing machine.
- Now it is universally accepted by computer scientists that TM is a Mathematical model of an algorithm.
- TM has an algorithm and an algorithm has a TM. If there is an algorithm problem is decidable, TMsolves that problem
- The statement of the thesis –

" Every function which would naturally be regarded as computable can be computed by a Turing machine"

Implies

- Any mechanical computation can be performed by a TM
- For every computable problem there is a TM
- If there is no TM that decides P there is no algorithm that can solve problem P.
- In our general life, we have several problems and some of these have solutions, but some have not, we simply say a problem is decidable if there is a solution otherwise undecidable.

example:

- Does Sun rises in the East? YES
- Will tomorrow be a rainy day ? (YES/NO ?)

Decidable and Undecidable Languages

- A problem is said to be decidable if its language is recursive OR it has solution.

Example:

Decidable :

-Does FSM accept regular language?

- is the power of NFA and DFA same

Undecidable:

- For a given CFG is L(G) ambiguous?

L is *Turing decidable* (or just decidable) if there exists a Turing machine Mthat accepts all strings in L and rejects all strings not in L. Note that by rejection means that the machine halts after a finite number of steps and announces that the input string is not acceptable.

- There are two types of TMs (based on halting):

1. **(*Recursive*)**
   TMs that *always* halt, no matter accepting or non-accepting ≡ **DECIDABLE** PROBLEMS
2. **(*Recursively enumerable*)**
   TMs that *are guaranteed to halt* only on acceptance.
   If non-accepting, it may or may not halt (i.e., could loop forever).
   - Undecidable problems are those that are not recursive

Recursive languages

A Language L over the alphabet∑ is called recursive if there is a TM M that accepts every word in L and rejects every word in L'

Accept (M)=L

Reject(M)=L' loop(M)=

ø Example: b(a+b)*

M is a *Turing Machine* and L is a *recursive language* that M accepts,
if a string w ∈ L then M *halts in a final state* and
if w ∉ L then M *halts in a non-final state*

Recursively Enumerable Language:

A Language L over the alphabet∑ is called recursively enumerable if there is a TM M that accepts every word in L and either rejects or loops every word in L' the complement of L

Accept (M)=L Reject(M)

+Loop(M)= L'

Example: (a+b)*bb(a+b)*

M is a *Turing Machine* and L is a *recursively enumerable language* that M accepts,
if a string w ∈ L then M *halts in a final state* and
if w ∉ L then M *halts in a non-final state or loops forever*

Theorem: If L is recursive language, so is $\overline{L}$

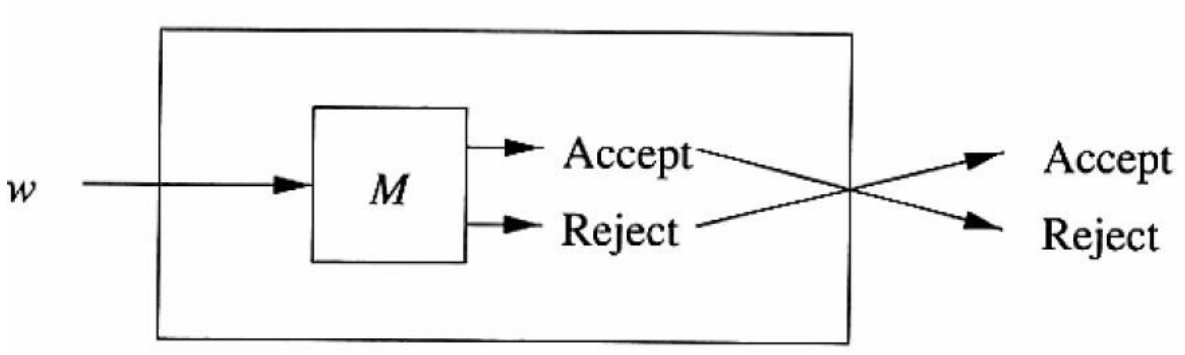
PROOF: Let $L = L(M)$ for some TM M that always halts. We construct a TM $\overline{M}$ such that $\overline{L} = L(\overline{M})$ by the construction suggested in Fig. 10 That is, $\overline{M}$ behaves just like M. However, M is modified as follows to create $\overline{M}$:

1. The accepting states of M are made nonaccepting states of $\overline{M}$ with no transitions; i.e., in these states $\overline{M}$ will halt without accepting.

2. $\overline{M}$ has a new accepting state r; there are no transitions from r.

3. For each combination of a nonaccepting state of M and a tape symbol of M such that M has no transition (i.e., M halts without accepting), add a transition to the accepting state r.

Since M is guaranteed to halt, we know that $\overline{M}$ is also guaranteed to halt. Moreover, $\overline{M}$ accepts exactly those strings that M does not accept. Thus $\overline{M}$ accepts $\overline{L}$. □



Recursively Enumerable Languages closed under complementation? (NO)

1. **Prove that Recursive Languagess are closed under Union**
   - Let M_u = TM for $L_1 \cup L_2$
   - M_u construction:
     1. Make 2-tapes and copy input w on both tapes
     2. Simulate M_1 on tape 1
     3. Simulate M_2 on tape 2
     4. If either M_1 or M_2 accepts, then M_u accepts
     5. Otherwise, M_u rejects.

2. **Prove that Recursive Languages are closed under Intersection**
   - Let M_n = TM for $L_1 \cap L_2$
   - M_n construction:
     1. Make 2-tapes and copy input w on both tapes
     2. Simulate M_1 on tape 1
     3. Simulate M_2 on tape 2
     4. If M_1 AND M_2 accepts, then M_n accepts
     5. Otherwise, M_n rejects.

3. Recursive languages are also closed under:
   a. Concatenation
   b. Kleene closure (star operator)
   c. Homomorphism, and inverse homomorphism
4. RE languages are closed under:
   a. Union, intersection, concatenation, Kleene closure
5. RE languages are *not* closed under:
   a. Complementation

1. **Decidable Languages about DFA : Prove that**

A_{DFA} is a decidable language.

$$A_{\text{DFA}} = \{\langle B, w\rangle \mid B \text{ is a DFA that accepts input string } w\}.$$

M = "On input $\langle B, w\rangle$, where B is a DFA and w is a string:

1. Simulate B on input w.
2. If the simulation ends in an accept state, *accept*. If it ends in a nonaccepting state, *reject*."

Proof: To prove we construct a TM that halts and also accept A_{DFA}
Define TM as

1. let B be a DFA and w input string (B,w) as input for TM M
2. Simulate B and input w in TM M
3. if the simulation ends in an accepting state of B then M accepts w. if it ends in non accepting state of B then M rejects w.

2. Prove that A_{NFA} is a decidable language.

$A_{NFA} = \{\langle B, w\rangle | \ B \text{ is an NFA that accepts input string } w\}.$

N = "On input $\langle B, w\rangle$ where B is an NFA, and w is a string:

1. Convert NFA B to an equivalent DFA C, using the **last** procedure .
2. Run TM M on input $\langle C, w\rangle$.
3. If M accepts, *accept*; otherwise, *reject*."

3. Prove that A_{REX} is a decidable language.

$A_{REX} = \{\langle R, w\rangle | \ R \text{ is a regular expression that generates string } w\}.$

P = "On input $\langle R, w\rangle$ where R is a regular expression and w is a string:

1. Convert regular expression R to an equivalent DFA A
2. Run TM N on input $\langle A, w\rangle$.
3. If N accepts, *accept*; if N rejects, *reject*."

Halting and Acceptance Problems:

$A_{TM} = \{\langle M, w\rangle | \ M \text{ is a TM and } M \text{ accepts } w\}.$

Acceptance Problem:

Does a Turing machine accept an input string?

1. A_{TM} is recursively enumerable.

Simulate M on w. if M enters an accepting state, We prove by contradiction. We assume **A_{TM} is decidable by a TM H that eventually halts on all input, then**

$$H(\langle M, w\rangle) = \begin{cases} accept & \text{if } M \text{ accepts } w \\ reject & \text{if } M \text{ does not accept } w. \end{cases}$$

$D =$ "On input $\langle M \rangle$, where M is a TM:

1. Run H on input $\langle M, \langle M \rangle\rangle$.
2. Output the opposite of what H outputs; that is, if H accepts, *reject* and if H rejects, *accept*."

$$D(\langle M \rangle) = \begin{cases} accept & \text{if } M \text{ does not accept } \langle M \rangle \\ reject & \text{if } M \text{ accepts } \langle M \rangle. \end{cases}$$

$$D(\langle D \rangle) = \begin{cases} accept & \text{if } D \text{ does not accept } \langle D \rangle \\ reject & \text{if } D \text{ accepts } \langle D \rangle. \end{cases}$$

We construct new TM D with H as a subroutine. D calls H to determine what M does when it receive the input <M>. Based on the received information on (M,<M>), D rejects M if M accepts <M> and accepts M if M rejects <M>.

D described as follows:
1. <M> is an input to D
2. D calls H to run on (M,<M>)
3. D rejects <M> if H accepts (M, <M>) and accepts <M>if H rejects (M, <M>)

This means D accepts <D> if D does not accept <D>, which is a contradiction. Hence **A_{TM}** is Undecidable.

The Post Correspondence Problem

PCP is a combinatorial problem formulated by Emil Post in 1946. This problem has many applications in the field theory of formal languages. A correspondence system P is a finite set of ordered pairs of non empty strings over some alphabet. Let $A = w_1, w_2, \ldots, w_n$ $B = v_1, v_2, \ldots, v_n$

There is a **P**ost **C**orrespondence Solution if there is a sequence $i, j, \ldots, k$ such that:

$$w_i w_j \cdots w_k = v_i v_j \cdots v_k$$

Index	w_j	V_i
1	100	001
2	11	111
3	111	11

Let W = $w_2w_1w_3 = v_2v_1v_3$ = 11100111 we have got a solution. But we may not get solution always for various other combinations and strings of different length. Hence PCP is undecidable.

The **M**odified **P**ost **C**orrespondence Problem

$A = w_1, w_2, \ldots, w_n$ $B = v_1, v_2, \ldots, v_n$

$1, i, j, \ldots, k$

$$w_1 w_i w_j \cdots w_k = v_1 v_i v_j \cdots v_k$$

If the index start with 1 and then any other sequence then it is called MPCP

Algorithm: An algorithm is "a finite set of precise instructions for performing a computation or for solving a problem"

- A program is one type of algorithm
  - All programs are algorithms
  - Not all algorithms are programs!
- The steps to compute roots of quadratic equation is an algorithm
- The steps to compute the cosine of 90° is an algorithm

Algorithms generally share a set of properties:

- Input: what the algorithm takes in as input
- Output: what the algorithm produces as output
- Definiteness: the steps are defined precisely
- Correctness: should produce the correct output
- Finiteness: the steps required should be finite
- Effectiveness: each step must be able to be performed in a finite amount of time

- Generality: the algorithm *should* be applicable to all problems of a similar form

Comparing Algorithms (While comparing two algorithm we use time and space complexities)

- Time complexity
  - The amount of time that an algorithm needs to run to completion
- Space complexity
  - The amount of memory an algorithm needs to run
- To analyze running time of the algorithm we use following cases
  - Best case
  - Worst case

    Average case

Asymptotic analysis

- The big-Oh notation is used widely to characterize running times and space bounds
- The big-Oh notation allows us to ignore constant factors and lower order terms and focus on the main components of a function which affect its growth
- Given functions $f(n)$ and $g(n)$, we say that $f(n)$ is $O(g(n))$ if there are positive constants c and n_0 such that

$$f(n) \leq cg(n) \text{ for } n \geq n_0$$

- Example: $2n + 10$ is $O(n)$
  - $2n + 10 \leq cn$
  - $(c - 2)\, n \geq 10$
  - $n \geq 10/(c - 2)$

  It is true for $c = 3$ and $n_0 = 10$

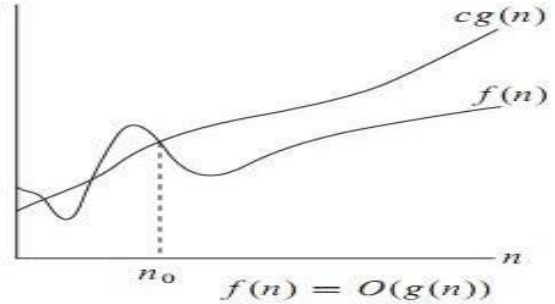
- $7n-2$ is $O(n)$

  need $c > 0$ and $n_0 \geq 1$ such that $7n-2 \leq c \bullet n$ for $n \geq n_0$

  this is true for $c = 7$ and $n_0 = 1$

f(n)= O(g(n)) *iff* there exist positive constants c and n0 such that f(n) ≤ cg(n) for all n ≥ n0

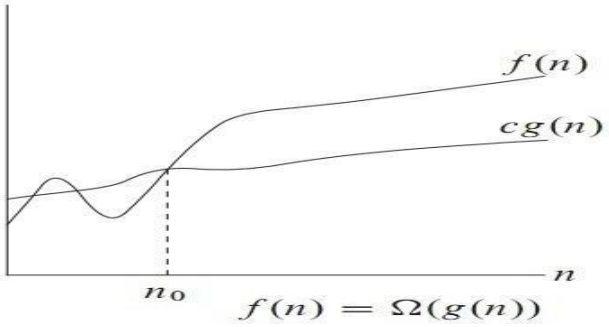
O-notation to give an upper bound on a function



Big oh provides an asymptotic upper bound on a function.

Omega provides an asymptotic lower bound on a function.

$$\Omega(g(n)) = \{f(n) : \text{there exist positive constants } c \text{ and } n_0 \text{ such that } 0 \le cg(n) \le f(n) \text{ for all } n \ge n_0\}.$$



- The big-Oh notation gives an upper bound on the growth rate of a function
- The statement "***f***(***n***) is ***O***(***g***(***n***))" means that the growth rate of ***f***(***n***) is no more than the growth rate of ***g***(***n***)
- We can use the big-Oh notation to rank functions according to their growth rate

$$f(n) = a_0 + a_1 n + a_2 n^2 + \ldots + a_d n^d$$

- If is ***f***(***n***) a polynomial of degree ***d***, then ***f***(***n***) is ***O***(n^d), i.e.,
  1. Drop lower-order terms
  2. Drop constant factors
- Use the smallest possible class of functions
  1. Say "2***n*** is ***O***(***n***)" instead of "2***n*** is ***O***(n^2)"
- Use the simplest expression of the class
  Say "3***n*** + 5 is ***O***(***n***)" instead of "3***n*** + 5 is ***O***(3***n***)"
- Following are the terms usually used in algorithm analysis:
  1. Constant ≈ ***1***
  2. Logarithmic ≈ log ***n***
  3. Linear ≈ ***n***
  4. N-Log-N ≈ ***n*** log ***n***
  5. Quadratic ≈ n^2
  6. Cubic ≈ n^3
  7. Exponential ≈ 2^n

Class P Problems:

P stands for deterministic polynomial time. A deterministic machine at each time executes an instruction. Depending on instruction, it then goes to next state which is unique. Hence time complexity of DTM is the maximum number of moves made by M is processing any input string of length n, taken over all input of length n.

- The class P consists of those problems that are solvable in polynomial time.
- More specifically, they are problems that can be solved in time $O(n^k)$ for some constant k, where n is the size of the input to the problem
- The key is that n is the **size of input**

Def: A language L is said to be in class P if there exists a DTM M such that M is of time complexity P(n) for some polynomial P and M accepts L.

Class NP Problems

Def: A language L is in class NP if there is a nondeterministic TM such that M is of time complexity P(n) for some polynomial P and M accepts L.

- **NP is not the same as non-polynomial complexity/running time. NP does not stand for not polynomial.**
- **NP = Non-Deterministic polynomial time**
- NP means verifiable in polynomial time
- Verifiable?
    - If we are somehow given a 'certificate' of a solution we can verify the legitimacy in polynomial time

- Problem is in NP iff it is decidable by some non deterministic Turing machine in polynomial time.
- It is provable that a Non Deterministic Turing Machine is equivalent to a Deterministic Turing Machine
- Remember NFA to DFA conversion?
    - Given an NFA with n states how many states does the equivalent DFA have?
    - Worst case …. 2^n
    - The deterministic version of a polynomial time
- non deterministic Turing machine will run in exponential time (worst case)
- Since it takes polynomial time to run the program, just run the program and get a solution
- But is NP a subset of P? It is not yet clear whether P = NP or not

Quantum Computers

- **Computers are physical objects, and computations are physical processes. What computers can or cannot compute is determined by the law of physics alone, and not by pure mathematics. Computation with coherent atomic-scale dynamics.**
  The behavior of a quantum computer is governed by the laws of quantum mechanics.

- In 1982 Richard Feynmann, a Nobel laurite in physics suggested to build computer based on uantum mechanics.
- Quantum mechanics arose in the early 1920s, when classical physics could not explain everything.
- QM will provide tools to fill up the gulf between the small and the relatively complex systems in physics.
- Bit (0 or 1) is the fundamental concept of classical computation and information. Classical computer built from electronic circuits containing wires and gates.
- Quantum bit and quantum circuits which are analogous to bits and circuits. Two possible states of a qubit (Dirac)are $|0\rangle$ $|1\rangle$
- Quantum bit is qubit described mathematically (where (alpha) is complex number)

$$\alpha_0|0\rangle+\alpha_1|1\rangle$$

- Qubit can be in infinite number of state other than dirac |0> or |1>

- he operations are induced by the apparatus *linearly*, that is, if

Then

$$\alpha_0|0\rangle+\alpha_1|1\rangle \rightarrow \alpha_0\left(\frac{i}{\sqrt{2}}|0\rangle+\frac{1}{\sqrt{2}}|1\rangle\right)+\alpha_1\left(\frac{1}{\sqrt{2}}|0\rangle+\frac{i}{\sqrt{2}}|1\rangle\right) = \left(\alpha_0\frac{i}{\sqrt{2}}+\alpha_1\frac{1}{\sqrt{2}}\right)|0\rangle+\left(\alpha_0\frac{1}{\sqrt{2}}+\alpha_1\frac{i}{\sqrt{2}}\right)|1\rangle$$

Any linear operation that takes states $\alpha_0|0\rangle+\alpha_1|1\rangle$ satisfying and maps them to be UNITARY

i.e. $|\alpha_0|^2+|\alpha_1|^2=1$

Linear Algebra: $|0\rangle$ Corresponds to $\begin{pmatrix}1\\0\end{pmatrix}$

$|1\rangle$ Corresponds to $\begin{pmatrix}0\\1\end{pmatrix}$

$\alpha_0|0\rangle+\alpha_1|1\rangle$ Corresponds to $\alpha_0\begin{pmatrix}1\\0\end{pmatrix}+\alpha_1\begin{pmatrix}0\\1\end{pmatrix}=\begin{pmatrix}\alpha_0\\\alpha_1\end{pmatrix}$

If we concatenate two qubits

$(\alpha_0|0\rangle+\alpha_1|1\rangle)$ $(\beta_0|0\rangle+\beta_1|1\rangle)$

we have a 2-qubit system with 4 basis states

$|0\rangle|0\rangle=|00\rangle$ $|0\rangle|1\rangle=|01\rangle$ $|1\rangle|0\rangle=|10\rangle$ $|1\rangle|1\rangle=|11\rangle$

And it is describes the state as $\alpha_0\beta_0|00\rangle+\alpha_0\beta_1|01\rangle+\alpha_1\beta_0|10\rangle+\alpha_1\beta_1|11\rangle$

- Quantum computer is a system built from quantum circuits, containing wires and elementary quantum gates, to carry out manipulation of quantum information.

Variants of Turing Machines

Various types of TM are

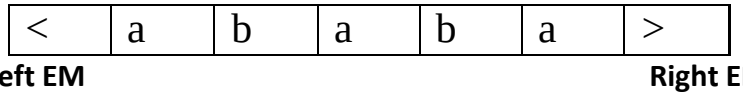
1. **With Multiple tapes**
2. **With one tape but multiple heads**
3. **With two dimensional tapes**
4. **Non deterministic TM**

1. **Multiple tapes: It consists of finite control with k tape heads and k tapes each tape is infinite in both directions. On a single move depending on the state of the finite control and symbol scanned by each of the tape head the machine can change state Or print new symbol on each of cell scanned etc..**

1. **With One tape but Multiple heads: a K head TM has fixed k number of heads and move of TM depends on the state and the symbol scanned by each head. (head can move left, right or stationary).**
2. **Multidimensional TM: It has finite control but the tape consists of a K-dimensional array of cells infinite in all 2 k directions. Depending on the state and symbol scanned , the device changes the state, prints a new symbol, and moves its tape head in one of the 2 k directions, either positively of negatively along one of the k axes.**
3. **Non deterministic TM: In the TM for a given state and tape symbol scanned by the tape head, the machine has a finite number of choices for the next move. Each choice consists of new state, a tape symbol to print and direction of head motion.**

Linear Bounded Automata

LBA is a restricted form of a Non deterministic Turing machine. It is a multitrack turing machine which has only one tape and this tape is exactly same length as that of input. It accepts the string in the similar manner as that of TM. For LBA halting means accepting. In LBA computation is restricted to an area bounded by length of the input. This is very much similar to programming environment where size of the variable is bounded by its data type. Lba is 7-tuple on Deterministic TM with



M= (Q,∑,Ƭ, Delta,q_{accept}, q_{reject}, q0)

- Two extra symbols < and > are used left end marker and right end marker.
- Input lies between these markers