Formal Languages and Compilers

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Contents

Course Information

Information about these notes

These notes are meant to be used in conjunction to lecture material provided from COSC261 at the University of Canterbury, there are references and links to particular lectures and videos about processes that will not be avalible to those reading this later

Course Staff

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Assessments and Grading

Grading policy

1. You mist achieve an average grade of at least 50% over all assessment items
2. You mist achieve an average grade of at least 45% over all invigilated assessments

Assessment Items

- Quizzes (15%)
- Assignment Superquiz (10%)
- Lab Test (15%)
- Final Exam (60%)

Textbooks/Resources

No textbooks are required, but see the following book for additional information:

- Carol Critchlow and David Eck; Foundationals of Computation; version 2.3.1, 2011

Lectures

Introduction

Topic overview of the course

- pattern matching
  - Regular expressions describe patterns
  - Search using REGEX is supported in many programs
  - Can all patterns be described by regular expressions?
  - One to one with state diagrams and automata
- Compilers
  - Programs can be run by an interpreter or by compiling them first
  - Interpreting may be slow
  - Compiling to machine code avoids much of the overhead
  - Compiler performs analysis, code generation and optimisation
  - How can these tasks be automated for different programming languages?
- Syntax analysis
  - Analyses code to determine if the syntax is correct for compiling, this is done by using context free grammers *there are other methods of doing this however this is the main one we will look at in this course*
  - *does the syntax conform to the languages grammar?*
  - Ideally we want to generate the parser for our language, we will look into how to manually like are parser and how regular expressions and pattern matching can be used to evaluate this behaviour.
- Code generation
  - There are formalisms that exist to generate code in order to create compilers via code generation.

Finite Automata and Regular Languages

Symbols, Strings and Languages

Languages

- An alphabet Σ is non-empty finite set of *symbols/*
- A string over Σ is a finite sequence of symbols from Σ
- The length of $|\sigma|$ of a string σ is the number of symbols in σ
- The empty string ϵ is the unique set of length 0
- Σ^* is the set of all strings over Σ
- A language L over Σ is a set of strings $L \subseteq \Sigma$

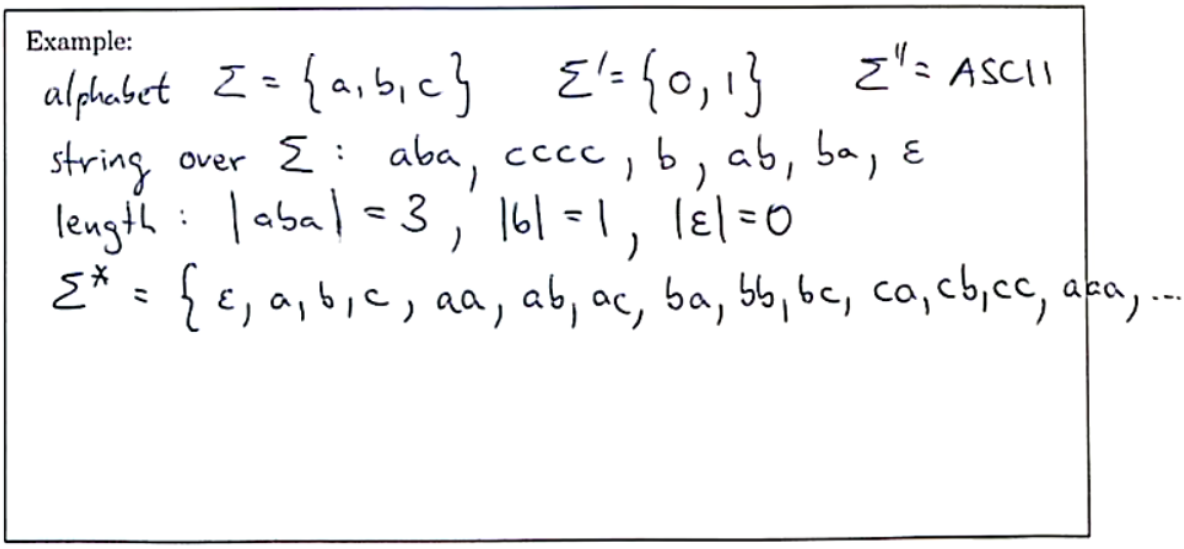


Figure 1: Example

- Note that with a finite alphabet we can have an infinite size for Σ
  - This is because we have not specified a size for our length of elements within Σ

> An example of a set that we might use is the unicode set as Sigma.

For example:

- $python \subset UNICODE$
- $english \subset UNICODE$

Because of this relationship, we can use filtering, searching and REGEX in order to manipulate and set rules around this relationship (or syntax in the case of programming languages by using comparisons and combination of formalisms.

Let $a, b \in \Sigma$ be symbols and let $x, y, z \in \Sigma$ be strings. - Symbols and strings can be concatenated by writing one after the other - xy is the concatenated version of x and y. - Note that concatenation is accociative - ϵ is an identity for concatenation $\epsilon x = x = x\epsilon$ - $|xy| = |x| + |y|$

Lifting to a set

$Let A, B \subseteq \Sigma$ be languages:

- concatenate languages A and B by conciliating each string from A with each from B
- $AB = \{xy | x \in A, y \in B\}$
- Language concatenation is associative
- $\{\epsilon\}$ is the identity of language concatenation
    - $x^0 = \epsilon$

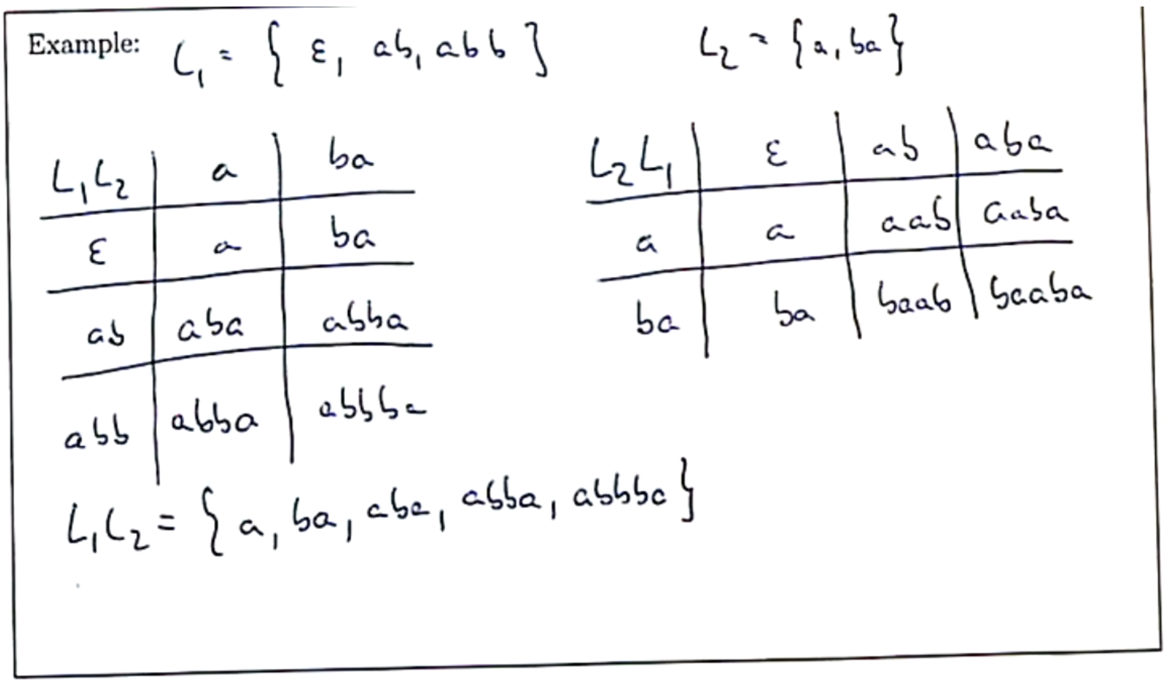


Figure 2: Lifting Example

Concatenation can be iterated

- a^n is the string comprising n copies of the symbol $a \in \Sigma$
- x^n is the string that concatinates n copies of the string $x \in \Sigma$

- These operations are defined *inductively*
- The base case is $x^0 = \epsilon$
- The *inductive case* is $x^{n+1} = x^n x$

Example: $a^5 = aaaaa$

A^n is defined similarly for a language $A \subseteq \Sigma^*$.

* $A^0 = \{\varepsilon\}$.
* $A^{n+1} = AA^n$.
* $A^1 = A,\ A^2 = AA,\ A^3 = AAA,\ \ldots$
* $A^* = \bigcup_{n\in\mathbb{N}} A^n = A^0 \cup A^1 \cup A^2 \cup A^3 \cup \ldots$
* $A^* = \{x_1x_2\ldots x_{n-1}x_n \mid x_i \in A \text{ for each } 1 \leq i \leq n \text{ for some } n \in \mathbb{N}\}$.
* $A^+ = \bigcup_{n\geq 1} A^n = AA^* = A^1 \cup A^2 \cup A^3 \cup \ldots$

Figure 3: Powers of Language

> Take aways, the * symbol means that we have zero or more of something, + means that we have one or more of something (this is how we use this notation practically in regex expressions

Key notation and definitions

- Sets are languages
- variables are strings
- variables with index are symbols

Deterministic Finite Automata

A *deterministic finite automaton (DFA)* is a structure $M = (Q, \Sigma, \delta, q_0, F)$ where:

- Q is a non-empty finite set, the *states*,
- Σ is a non-empty finite set, the *input alphabet*
- $\delta : Q \times \Sigma \to Q$ is the *transition funtion*
- $q_0 \in Q$ is the *start state*,
- $F \subseteq Q$ is the set of *accept states* or *final states*.

This can be shown in a *transition diagram* for a visual indication of how this might work:

> see lecture 3, 14:00 minutes for information on how to construct these transition diagrams

Deterministic Finite Automata Example Ex 1: The following DFA accepts strings over $\{a, b\}$.

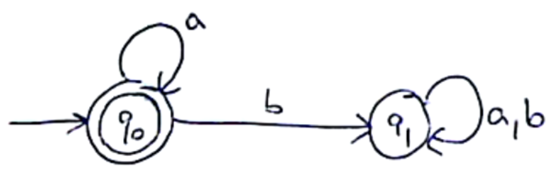


Figure 4: Sink state diagram

Note, we are trying to conatain an automata that is as least complicated as possible (least nodes). In the above example, we have this concept of a true state and a false state to the specified condition. We can also note that state q_1 is a *sink state* as once we reach q_1 it is impossible to leave that state (*this is because the condition specifies that we need to contain 0 b's*).

Ex 2: DFA accepting all strings over $\{a, b\}$ with a number of a-symbols that is not a multiple of 4.

Closure of properties of Deterministic Finite Automata **Exptended transition function**

$$\hat{\delta} : Q \times \Sigma \to Q$$

$$\hat{\delta}(q, \epsilon) = q$$

$$\hat{\delta}(q, ax) = \hat{\delta}(\delta(q, a), x) \quad where \quad a \in \Sigma, x \in \Sigma$$

Now we have seen the basics of DFA's, so how do we get from these basic models to something that is more complicated? We can use a combination of DFA's in order to define more complicated systems.

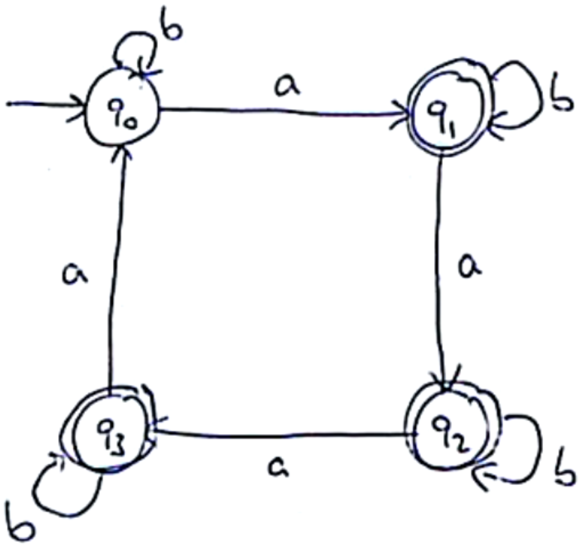


Figure 5: Modulo four example

Regular languages are closed under:

- complement
  - All strings that are not in the set, (generic approach to create such an atomaton)
  - We can get the complement by swapping accepting and non-accepting states
  - If we have a regular language, we know that the complement will always be regular
    * Formal proof provided in `Lecture Four: 27:10`
- intersection *or product automaton*
  - We can combine two automata using an *intersection*, we can use set theory in order to satisfy two automata at the same time. We want to check with one automataon to see if the string is satisfied
  - $X \cap Y$ is regular if both X and Y are regular
  - Accept states are defined as an acceptance of **both** automata, not just one
  - The product automaton accepting the intersection of the two languages is (*synchronous*):
    * Example found in `Lecture Four: 41:00`
- union
- concatenation
- star

Non-Deterministic Finite Automata

The following automaton accepts strings with a symbol 1 in the third position from the end. It is not a DFA because there are two 1-transitions in state q_0 and no transitions in state q_3.

- **DFA defined by**: $\delta : Q \times \Sigma \to Q$
- **NFA defined by**: $\delta : Q \times \Sigma \to P(Q)$
  - where $P(Q) = \{S | S \subseteq Q\}$, $P(Q)$ is also called the *power set*

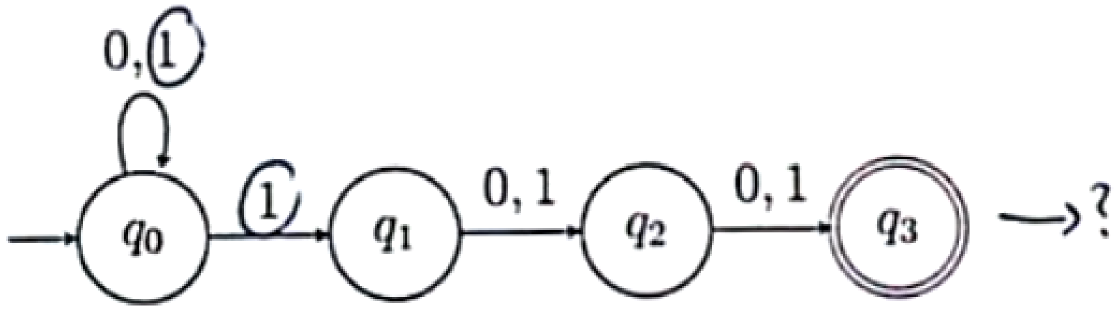


Figure 6: Non-Deterministic Finite Automata

Here is an example of the above transition relation:

δ	0	1
q_0	$\{q_0\}$	$\{q_0, q_1\}$
q_1	$\{q_2\}$	$\{q_2\}$
q_2	$\{q_3\}$	$\{q_3\}$
q_3	$\{\emptyset\}$	$\{\emptyset\}$

The extended transition relation $\hat{\delta} : Q \times \Sigma^* \to P(Q)$ is

- $\hat{\delta}(q, \epsilon) = \{q\}$
- $\hat{\delta}(q, ax) = \bigcup_{p \in \delta(q,a) \hat{\delta}(p,x)}$ where $a \in \Sigma$ and $x \in \Sigma^*$

Example of evaluating this extention can be found in `Friday March 4th Lecture: 8:00`

> The result of this will tell you what states are avalible from the current state (in a recursive nature)

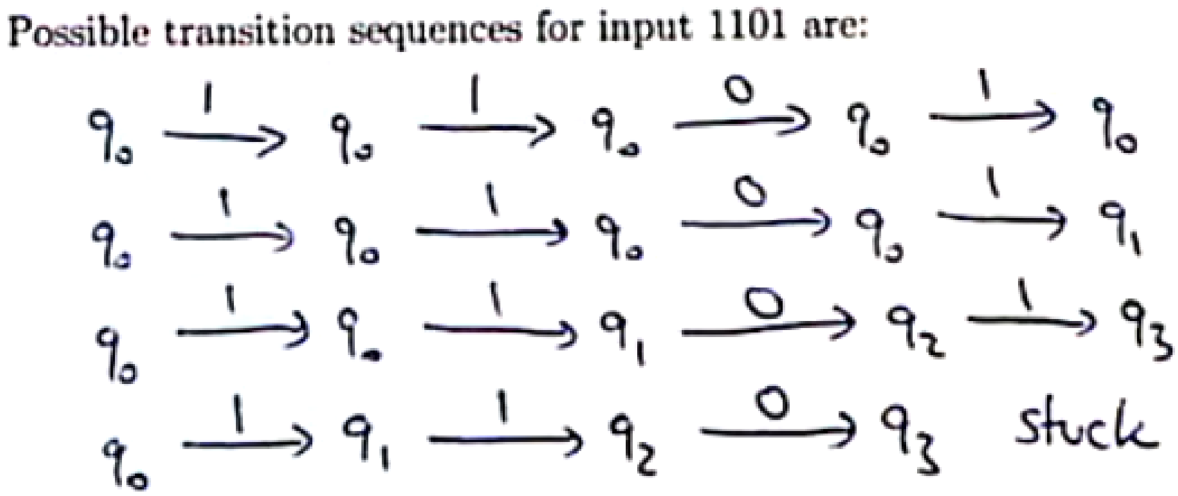


Figure 7: Possible transition states of this non-deterministic finite automata

> If we check the acceptance criteria for the automaton, we can do this by using this method to get to the last state.

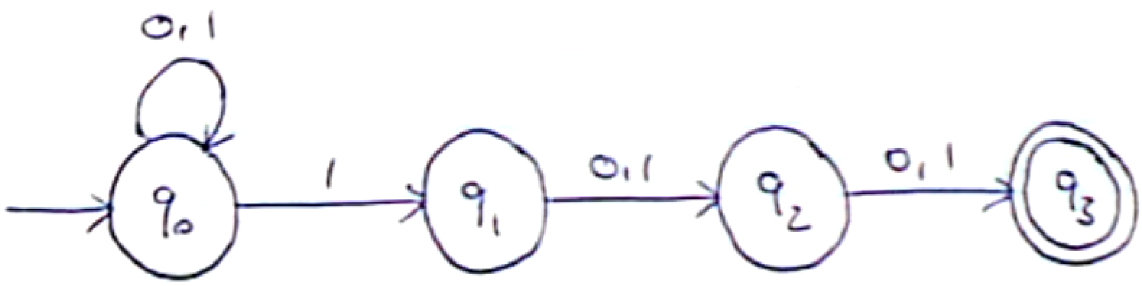


Figure 8: Consider this Automaton

The Subset Construction We need to consider all the possible nodes that we can reach from the current state. Below is an example of how we might do such a traverse on the above automaton. Note that we don't know what state we are going to be in, we are considering all possible options at a point in time.

In the following example, the numbers in each node are shorthand notation for the set containing all possible nodes that we can reach from that value, hence $q_{012} \to \{q_0, q_1, q_2\}$.

The above construction considers all possibilities of the `non-deterministic automaton`, this allows us to construct a `deterministic finite automata`, the only change we need to make

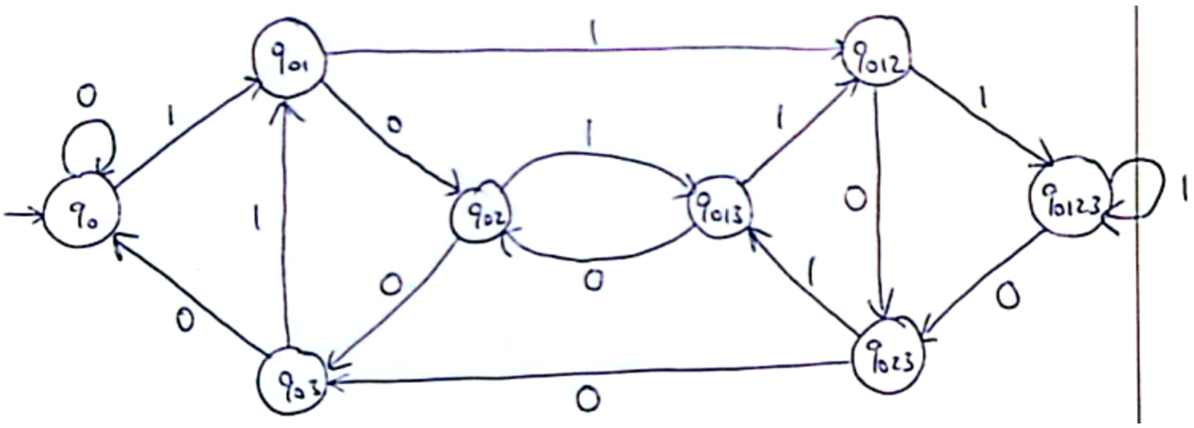


Figure 9: Subset Construction

is to add accpetance states to this construction.

Every language accepted by an NFA is accepted by a DFA

Proof: NFA $M = (Q, \Sigma, \delta, q_0, F)$

idea: keep track of all states M can be in while reading the input

subset automaton (DFA) $M' = (\mathcal{P}(Q), \Sigma, \delta', \{q_0\}, F')$

$$F' = \{ S \subseteq Q \mid S \cap F \neq \emptyset \}$$

$$\delta'(S, a) = \bigcup_{q \in S} \delta(q, a)$$

$$\Rightarrow \hat{\delta}'(S, w) = \bigcup_{q \in S} \hat{\delta}(q, w) \quad \circledast$$

To show $L(M') = L(M)$. For every $w \in \Sigma^*$,

$$
\begin{aligned}
& w \in L(M') \\
\overset{\text{Def. DFA } L(M')}{\iff}\ & \hat{\delta}'(\{q_0\}, w) \in F' \\
\overset{\text{Def. } F'}{\iff}\ & \hat{\delta}'(\{q_0\}, w) \cap F \neq \emptyset \\
\overset{(*)}{\iff}\ & \Big(\bigcup_{q \in \{q_0\}} \hat{\delta}(q, w)\Big) \cap F \neq \emptyset \\
\overset{\text{Def. } \cup}{\iff}\ & \hat{\delta}(q_0, w) \cap F \neq \emptyset \\
\overset{\text{Def. NFA } L(M)}{\iff}\ & w \in L(M)
\end{aligned}
$$

> Through this lecture, we have found an equivilence relation between DFA's and NFA's, meaning that we can convert DFA's to NFA's, therefore NFA's accept exactly regular languages. The number of states may grow exponentially in the subset construction.

Non-Deterministic Finite Automata with ϵ-Transitions THe following automaton accepts the union of two regular languages. It is not a DFA becuase of the ϵ-transitions in state r_0.

An *NFA* with ϵ transitions is a structure $M = (Q, \Sigma, \delta, q_0, F)$ where:

- Q, Σ, q_0 and F are as in a *DFA*,
- ϵ is a special symbol with $\epsilon \notin \Sigma$,
- $\delta : Q \times (\Sigma \cup \{\epsilon\}) \to P(Q)$ is the transition relation
- δ may have ϵ-transitions and yields a set of successor states

This is a way to decouple and clarify a set of states and a choice of two transitions in an automaton, this allows us to clearly identify and mark unions in the automaton.

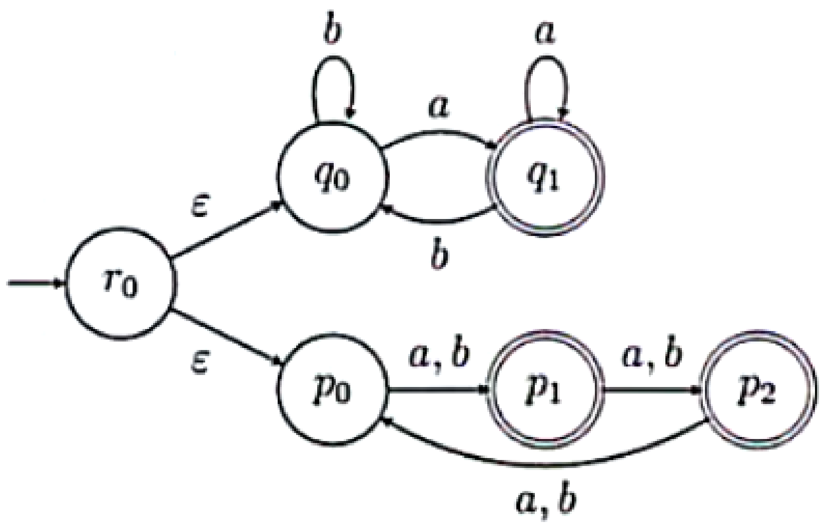


Figure 10: Epsilon transitions

Here is an example of a *NFA* that is designed in order to check validity of different types of numbers for a programming language. *Hexadecimal, decimal, octal and binary*.

Regular Expressions

These are partterns that can be used to match substrings in a given string:

- `ls *201?.*` lists files whose name without extension ends in 201 followed by some character
- `ls *a*a*a*` lists any file whose name contains three `a` charaters
- `rm *.log` deletes all log files
- `grep '[A-Z][a-z]\\{3,7\\}'` finds lines with a capital followed by 3-7 lower case letters

Atomic patterns are:

- $a \forall a \in \Sigma$ is matched by the symbol a
- ϵ is matched by the empty string
- θ is matched by nothing
- ? is matched by any symbol in Σ

Compound patterns are formed from patterns *p* and *q* as follows:

- $p|q$ is matched by string w if w matches p or q.
- pq is matched by w if $w = xy$ and x matches p and y matches q.
- $\not p$ matches by w if if w does not match p.

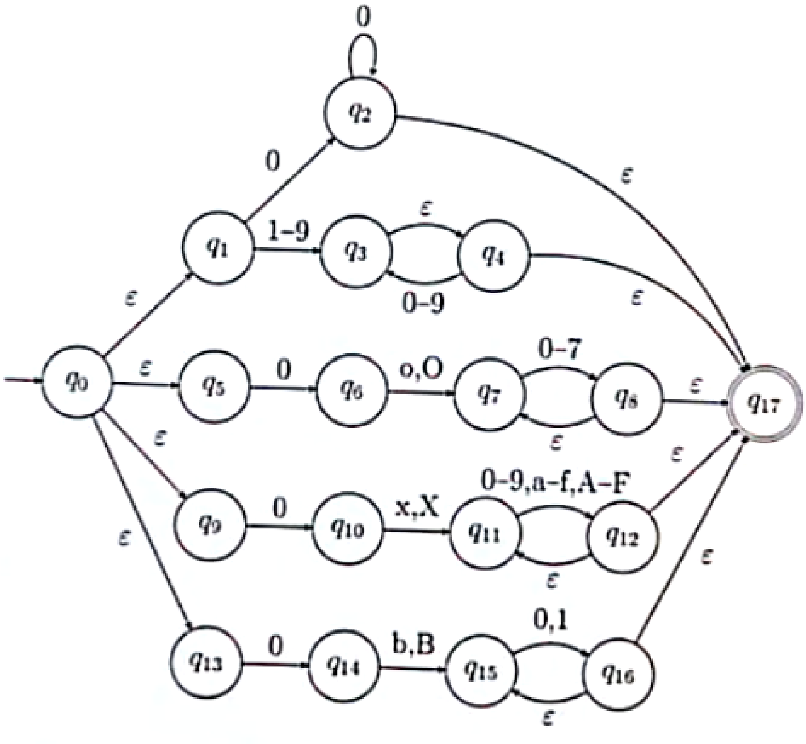


Figure 11: numerical automaton

- $[p]$ is matched by w if w is empty or matches p.

Here are some examples of a compound expression that can be evaluated in order to get its set of mapped values.

$$L((a|b)) = L(a) \cup L(b) = \{a\} \cup \{b\} = \{a, b\}$$

$$L((a|b)(c|a)) = L(a) \cup L(b)L(c) \cup L(a) = \{a, b\} \cup \{c, a\} = \{ac, aa, bc, ba\}$$

In order to preserve readability, we can ommit parentheses in order to simplify the expressions. Here is an example: $p|qr* = (p|(q(r)*))$

It is important to realise that we can construct ANY regular expression, and all such expressions will be accepted by a corresponding NFA.

Creating Automata to only accept single constructs

For ϵ construct:

```
     --    E     --
-->|   | ---> |   |
     --          --
```

For $\emptyset$ construct:

```
     --          --
-->|   |        |   |
     --          --
```

More examples can be found in `Lecture 8: 30:00`

> NOTE: include all standard constructs for automatons that are commonly used

The above automatons can be used for formulation of more complicated regular expressions, we simplify these **at your own risk** as we may be able to save some states (making program more efficent) however we may find that we are leaving out cases that should be covered by our automaton.

Using Regular Expressions as Symbols within Automata

- Transitions may be labelled with regular expressions instead of just symbols from Σ or ϵ
- There is just one accept state
- There is at most one transition between any two states
- There are no transtions into the start state or out of the acept state

We will see how we can translate any automaton into a regular expression. The main idea behind how we achieve this is to try to eliminate states from the diagram.

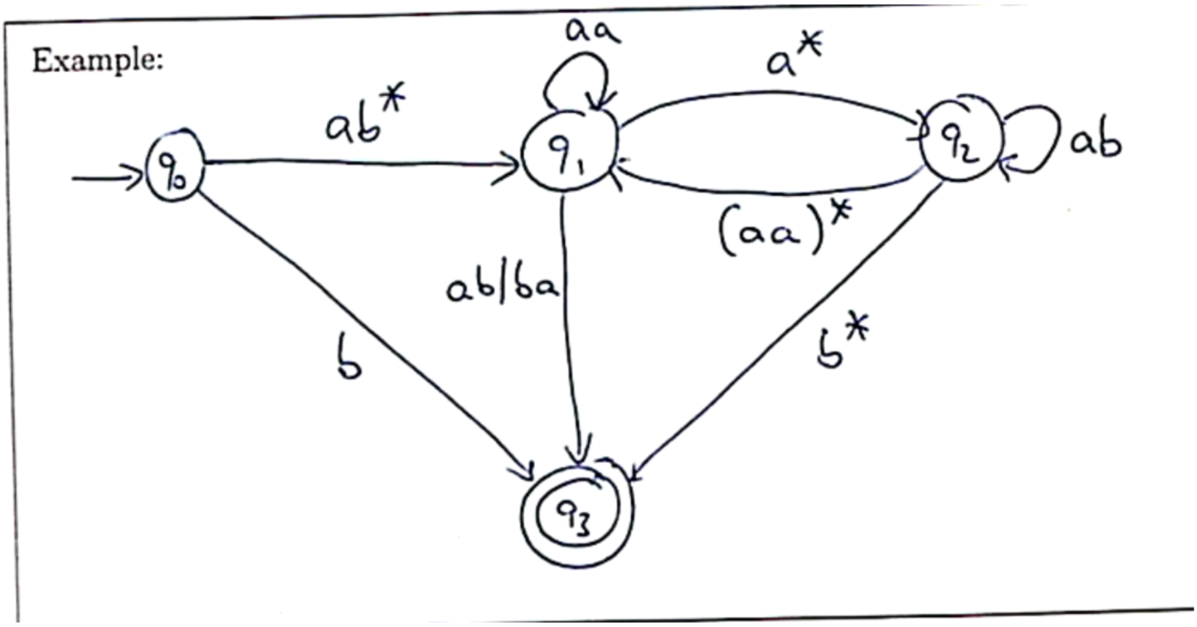


Figure 12: Regex Transitions Example

Every language accepted by an NFA is generated by a regular expression. Proof:

- Idea: successively eliminate states from generalised NFA.
- Done if just the start state and the accept state are left.
- Otherwise choose a state q_j that is neither the start state nor the accept state.
- Make the following change for all states q_i and q_k with $i \neq j$ and $k \neq j$:

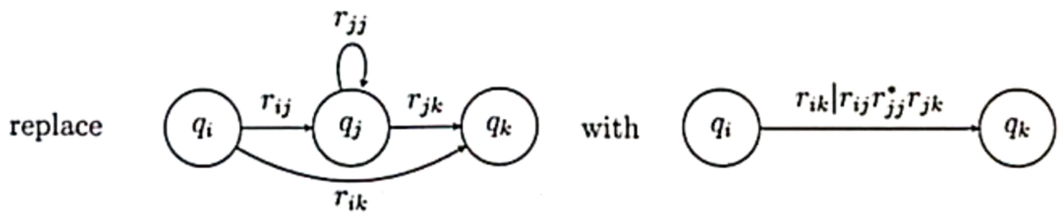


Figure 13: Converting automaton to regex

> The hard part of the above process is that we must apply this algorithm to all possible nodes and edges, this means that we must apply the same pattern to all possible ways (this is unusual and some times difficult).

Minimisation of Deterministic Finite Automata

A method to minimise the number of states in a DFA is:

1. Eliminate states which cannot be reached from the start state
2. Find equivalent states
3. Collapse equivilent ststes

The two states are *equivilant* if:

- $\hat{\delta}(p, w) \in F \iff \hat{\delta}(p, w) \in F \forall w \in \Sigma^*$
- The automation accepts the same strings when started in *p* or in *q*.
- Collapsing *p* and *q* does not change the accepted language
- *p* and *q* are *distinguishable* if they are not equivalent

Distringuishable states can be obtained as follows:

- Any $p \in F$ and $q \notin F$ are distinguishable by $w = \epsilon$
- Let $\delta(p, a) = r$ and $\delta(q, a) = s$ for $a \in \Sigma$
    - If r and s are distinguishable by ($w = x$) then p and q are distinguishable by ($w = ax$)

Equivilence in finite automata

We use the notation `p ~ q` if states `p` and `q` are equivilant, the relation ~ has:

- is reflexive: $p \sim p \quad \forall \quad p \in Q$
- is symmetric: $p \sim q$ implies $q \sim p$ for all $p, q \in Q$
- is transitive $p \sim q$ and $q \sim r$ implies $p \sim r$ for each $p, q, r \in Q$

An *equivilence relation* is a relation $\sim \subseteq A \times A$ that is reflexive, symmetric and transitive.

- $[a] = \{b \in A | a \sim b\}$ is the equivalence class of $a \in A$
- a is representative of its equivilence class `[a]`
- $A/\sim = \{[a] | a \in A\}$ is the *quotent* of A by $\sim$
    - This is the set of all equivialence classes

Automaton can start from any nodes that are found in the equivilence class that starting node s_0 is in.

Note that we minimal DFA's are unique up to isomprphism

The minimisation algorithm makes a quotient construction:

* Let DFA $M = (Q, \Sigma, \delta, q_0, F)$ be given.
* The *quotient automaton* is $M' = (\underbrace{Q/\sim}, \Sigma, \delta', [q_0], F')$.
* A state of M' is an equivalence class of $\sim$.
* $F' = \{[q] \mid q \in F\}$ comprises the equivalence classes of the accept states of M.
* $\delta' : \underbrace{Q/\sim} \times \Sigma \to \underbrace{Q/\sim}$ is defined by $\delta'([q], a) = [\delta(q, a)]$.
* δ' is well-defined because $[\delta(q, a)]$ does not depend on which representative q is chosen.

Figure 14: Algorithm to minimise quotient

Minimisation algorithm Quotient construction

Decision Problems for Regular Languages Simple yes or no problems about the properties of a given set

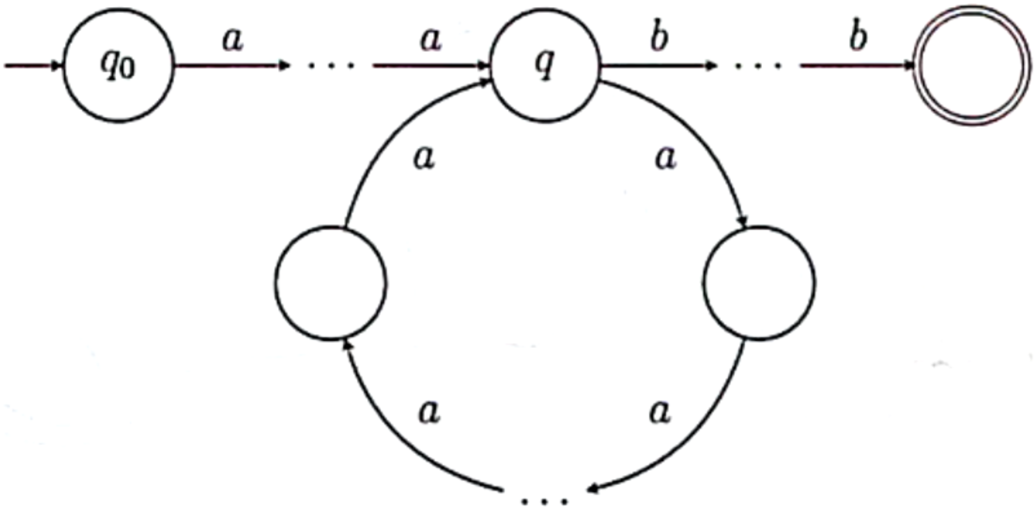
> See Lecture 10, 24:00 minutes to see how these are done Note, Go through these examples and actually implement them for end of year exams [ ]

Non-Regular Languages

The language $A = a^n b^n | n \in \mathbb{N}$ is a non-regular language.

The idea behind this proof is that we can construct a cycle where some iterations of the cycle are not included in the acceptance state.

- Assume that A is regular, we can prove this by applying contradiction
  - $A = L(M)$ for DFA $M = (Q, \Sigma, \delta, q_0, F)$ with k states
- The transition sequence for input a^k contains $k + 1$ states
- By the pigeonhole principle, a state q is visited twice

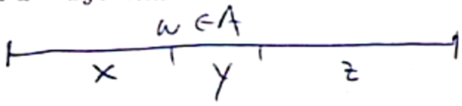


- There are i, j with $- \leq i \leq j \leq k$ and $\hat{\delta}(q_0, a^i) = \hat{\delta}(q_0, a^j)$
- Hence $\hat{\delta}(q_0, a^i b^i) = \hat{\delta}(q_0, a^j b^i)$
- But $a^i b^i \in A$ and $a^j b^i \notin A$, so $a^i b^i \in F$ and $a^j b^i \notin F$
- This is a contradiction, so the assumption does not hold

The Pumping Lemma The pumping lemma can be used to identify a contradiction to be used in a proof.

Pumping Lemma: Let A be regular. Then there is a number n such that each $w \in A$ with $|w| \geq n$ can be broken into three parts $w = xyz$ with

* $y \neq \varepsilon$,
* $|xy| \leq n$, and
* $xy^iz \in A$ for each $i \in \mathbb{N}$.



$xz \in A$
$xyz \in A$
$xy^2z \in A$
$xy^3z \in A$
$\vdots$

Use the pumping lemma to show that $A = \{a^nb^n \mid n \in \mathbb{N}\}$ is not regular:

Assume A is regular.

Let $n \in \mathbb{N}$.

Set $w = a^nb^n$.

Let x, y, z be such that $w = xyz$ and $y \neq \varepsilon$ and $|xy| \leq n$. $\rightsquigarrow$ $y = a^j$ for $j \geq 1$

Set $i = 2$.

Then $xy^iz \in A$ by the pumping lemma, but $xy^2z = xyyz = a^{n+j}b^n \notin A$. contradiction.

The pumping lemma is an exchange of information between a set of agents, we are able to obtain certain values from this lemma in order to obtain more information about the problem we are trying to prove.

Sorting a set

$\forall$ input : input is a set $\Rightarrow$
$\exists$ output : set(output) = input $\wedge$
len(output) = |input| $\wedge$
$\forall i$: $0 < i <$ len(output) $\Rightarrow$
output[i-1] $\leq$ output[i]

Figure 15: Sorting a list example

The above figure is displaying how we might solve a programming problem using mathematics and the pumping lemma, the take away of this is that when we are solving problems in computer science, we are essentailly proving a mathematical statement. If we can prove the method from this construction, we do not need to test the output, as we can guarentee that if the proof holds and is the same as the problem construction, then the program has to work with 100% consistancy.

The pumping lemma shows us that a few languages are non-regular, and if we can map other languages to the pumping lemma's non-regular languages, then we can determine that that language is also non-regular

Modelling Independent Processes

This is a model of concurrency within a program. The *shuffle* operation describes two independent processes:

- The shuffle `x || y` if two strings $x, y \in \Sigma^*$ contains all possible interleavings of their symbols
- For example, `ab||cd = {abcd, acbd, acdb, cabd, cadb, cdab}`
- Shuffle of strings is defined inductively:

$$
\begin{aligned}
\varepsilon \| y &= \{y\} \\
x \| \varepsilon &= \{x\} \\
ax \| by &= \{a\}(x \| by) \cup \{b\}(ax \| y) \qquad \text{for } a, b \in \Sigma
\end{aligned}
$$

Figure 16: Shuffle of strings

Context-Free Languages

Context-Free Grammars

A *context-free grammar (CFG)* is a structure $G = (N, \Sigma, P, S)$ where

- N is a finite set, the *non-terminals*
- Σ is a finite set disjoint from N, the *terminals*
- $P \subseteq N \times (N \cup \Sigma)^*$ is a finite set of *Productions*
- $S \in N$ is the *start symbol*

Productions are denoted as follows:

- A production $(A, w) \in P$ is written $A \to w$
- Several productions are written as $A \to w_1|...|w_n$
- THe right-hand side may be empty: an ϵ-production is written $A \to \epsilon$

Here is an example of G that matches with arithmetic expressions:

$$\Sigma = \{+, *, (,), n\}$$

$$N = \{E, T, F\}$$

$$S = E$$

$$P = E \to T|E + T$$

$$T \to F|T * F$$

$$F \to (E)|n$$

Note that the production notation can be expressed as a piecewise

$$E = \begin{cases} E \to T \\ E \to E + T \end{cases}$$

This is the same expression as shown above in the example

Using the previous example, we will see how we can compare strings using these CFG's, see Lecture: `Tue Mar 22 : 16 Minutes`.

Chomsky Normal Form A CFG $G = (N, \Sigma, P, S)$ is in chomsky normal form if every produciton has the form:

- $A \to BC$ where $B, C \in N$ or,
- $A \to a$ where $a \in \Sigma$
- The right hand side of each production is either two non-terminals or a terminal

An example of how to convert to Chomsky Normal Form is run through in `Fri Mar 25 Lecture: 20 Minutes`

The Cocke-Younger-Kasami Algorithm This is a recursive algorithm to solve the membership problem $w \in L(G)$, we start by assuming that G is in chomsky normal form, then we let n be the length of

For every CFG G with $\varepsilon \notin L(G)$ there is a CFG G' in Chomsky normal form with $L(G) = L(G')$.

1. Eliminate ε-productions of the form $A \to \varepsilon$.
2. Eliminate unit-productions of the form $A \to B$.
3. Eliminate non-generating non-terminals.
4. Eliminate non-reachable non-terminals.
5. Eliminate terminals from right-hand sides of length at least 2.
6. Eliminate right-hand sides of length at least 3.

1. If $A \to uBv \in P$ for $u, v \in (\Sigma \cup N)^*$ and $B \to \varepsilon \in P$, add $A \to uv$ to P.

$*$ Repeat this step while there are changes.

$*$ Afterwards, remove all ε-productions of the form $A \to \varepsilon$.

2. If $A \to B \in P$ and $B \to w \in P$ for $w \in (\Sigma \cup N)^*$, add $A \to w$ to P.

$*$ Repeat this step while there are changes.

$*$ Afterwards, remove all unit-productions of the form $A \to B$.

3. A non-terminal A is *generating* if $A \Rightarrow^* w$ for some $w \in \Sigma^*$.

$*$ If $A \to w \in P$ and w contains only terminals, A is generating.

$*$ If $A \to w \in P$ and w contains only terminals or generating non-terminals, A is generating.

$*$ Remove each non-generating non-terminal with all productions containing it.

4. A non-terminal A is *reachable* if $S \Rightarrow^* uAv$ for some $u, v \in (\Sigma \cup N)^*$.

$*$ S is reachable.

$*$ If $A \to w \in P$ and A is reachable, every non-terminal in w is reachable.

$*$ Remove each non-reachable non-terminal with all productions containing it.

5. Consider every terminal a in a right-hand side of length at least 2.

$*$ Add a new non-terminal A and the production $A \to a$.

$*$ Replace every occurrence of a in a right-hand side of length at least 2 with A.

6. Consider every production $A \to B_1 B_2 \ldots B_n$ with $n \geq 3$.

$*$ Add a new non-terminal C and replace this production with $A \to B_1 C$ and $C \to B_2 B_3 \ldots B_n$.

$*$ Repeat this step while there are changes.

Figure 17: Steps for converting to Chomsky Normal Form

w. Mark the indicies for the symbols in w and then computer between positions *i,j* until we generate w_{ij}. The end result is calculating $N_{ij} = \{A \in N | A \longrightarrow^* w_{ij}\}$.

Given a string $w \in \Sigma^*$ and a CFL A, is $w \in A$?

- This is the test for *membership* in a CFL.
- It is similar to parsing, but no syntax tree has to be constructed.
- $A = L(G) = \{x \in \Sigma^* \mid S \Rightarrow_G^* x\}$ for a CFG G.
- Hence $w \in A$ if and only if $S \Rightarrow_G^* w$.
- Checking all derivations does not work, since there might be infinitely many.
- Assume that G is in Chomsky normal form.
- Every non-terminal produces at least one terminal.
- It suffices to consider derivations that introduce up to $|w|$ non-terminals.
- This gives an upper bound on the length of derivations that need to be checked.
- The number of derivations might still be exponential in the length of w.

This is a technique to improve running time, it is a specific use of *dynamic programming* (bottom up approach). It works by creating smaller problems in order to solve bigger problems and avoids repeating calculation of previous smaller problems.