Computational Models - Lecture 5[1]

Handout Mode

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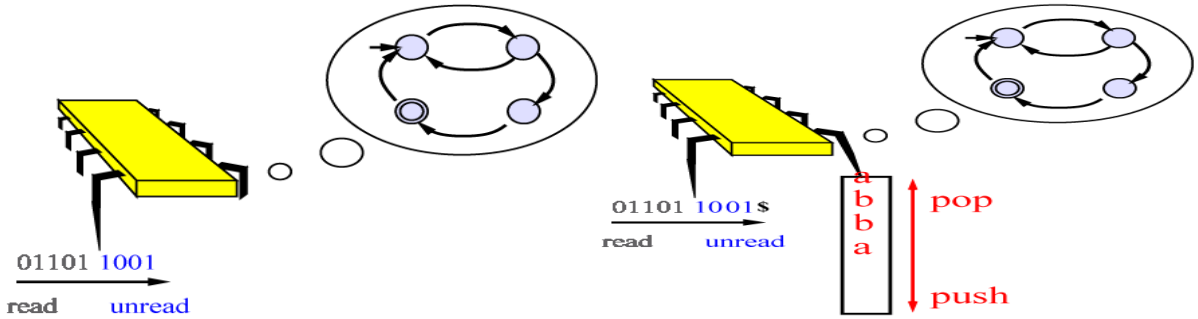
April 10/22, 2013

[1] Based on frames by Benny Chor, Tel Aviv University, modifying frames by Maurice Herlihy, Brown University.

Outline

- Push Down Automata (PDA)
- Equivalence of CFGs and PDAs
- Closure properties for CFL



- Sipser's book, 2.1, 2.2 & 2.3

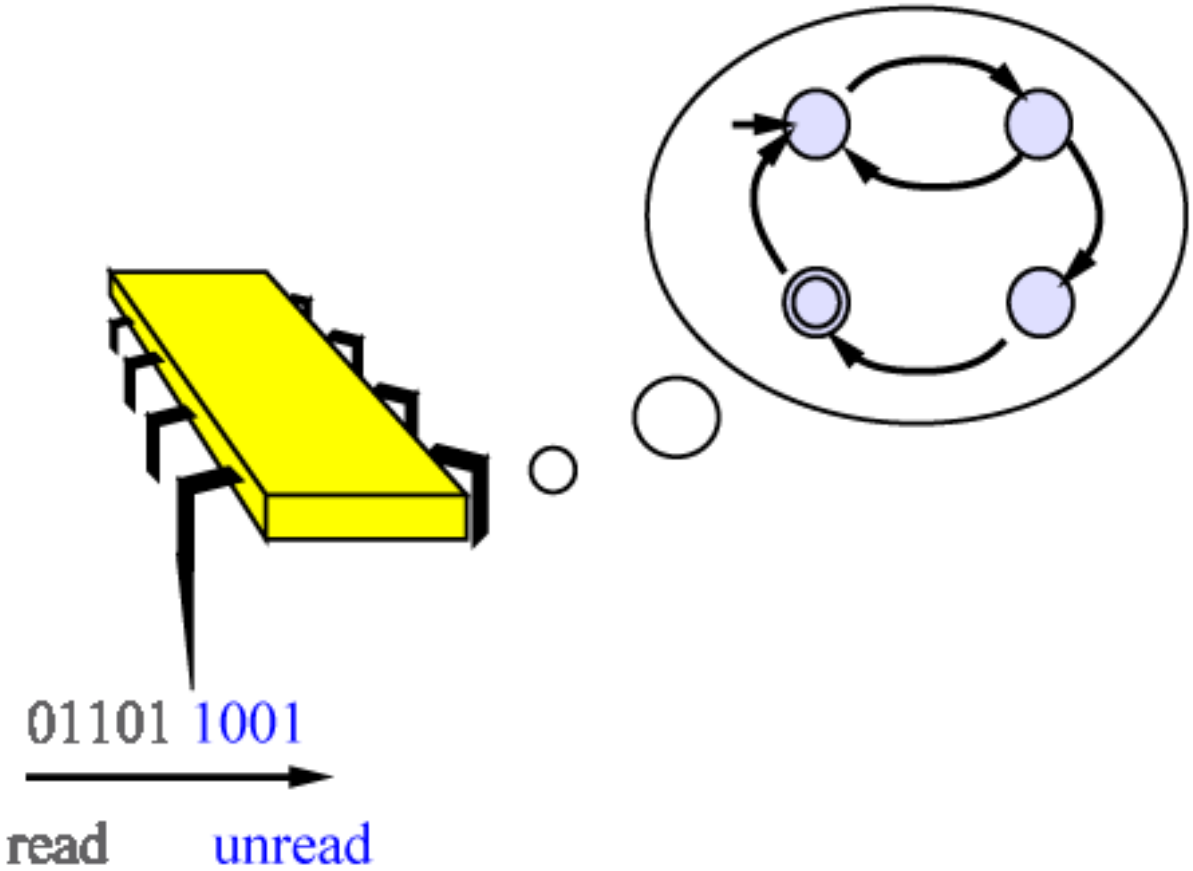
Part I

Push-Down Automata

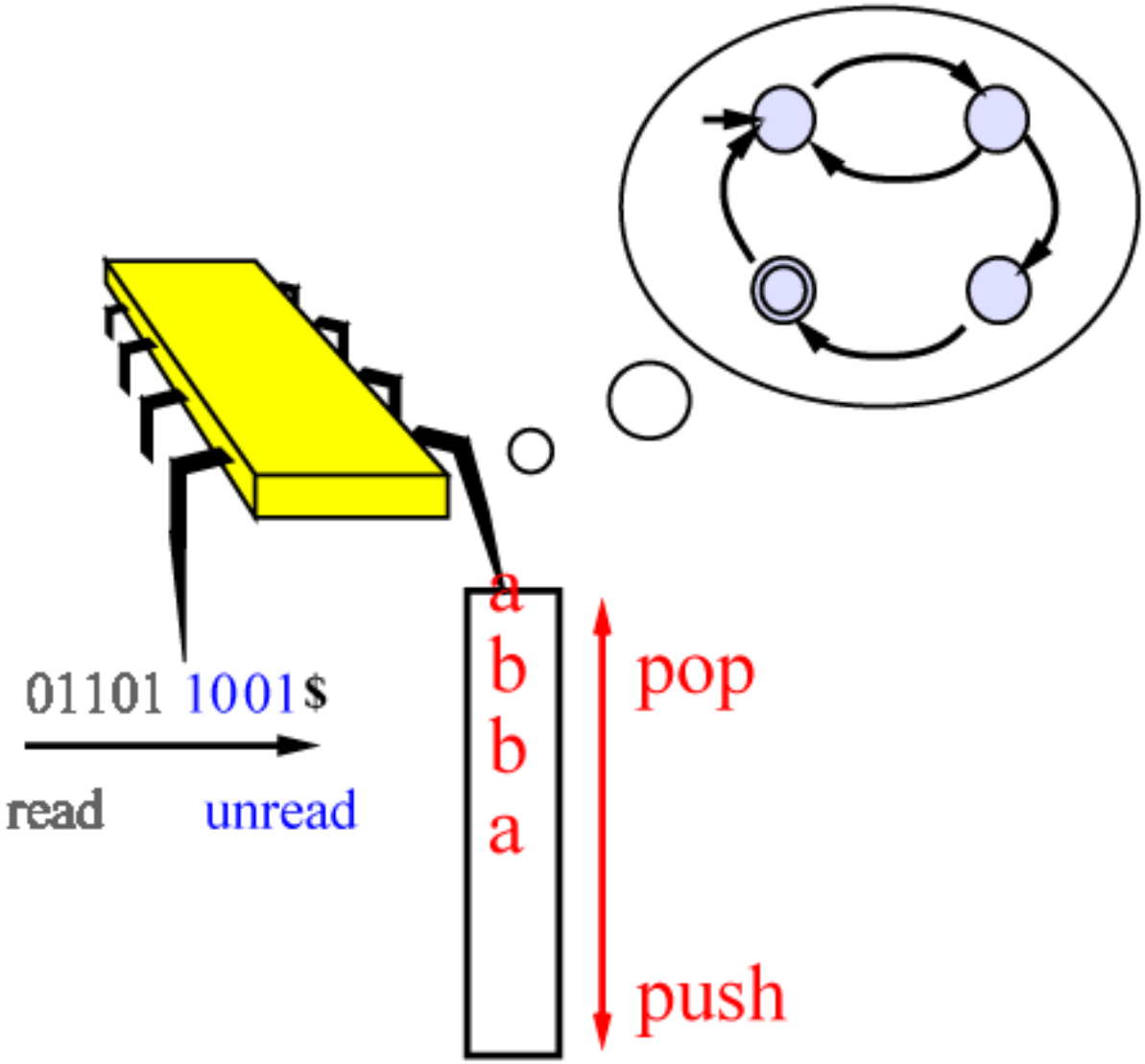
String Generators and String Acceptors

- Regular expressions are string generators – they tell us how to generate all strings in a language $\mathcal{L}$
- Finite Automata (DFA, NFA) are string acceptors – they tell us if a specific string w is in $\mathcal{L}$
- CFGs are string generators
- Are there string acceptors for CFLs?
- YES! *Push-down automata*

A Finite Automaton



A PushDown Automaton



(ignore the '$' sign)

Example 1 — PDA for $\mathcal{L}_1 = \{0^n 1^n : n \geq 0\}$

Informally:

1. Read input symbols
   1. Push each read 0 on the stack
   2. Pop a 0 for each read 1
2. `Accept` if stack is empty after last symbol read, and no 0 appears after 1

Recall that $\mathcal{L}_1$ is not regular

Example 2 — PDA for $\mathcal{L}_2 = \{a^i b^j c^k : i = j \vee i = k\}$

Informally:

Read and push a's
Either pop and match with b's or pop and match with c's

A non-deterministic choice

Pushdown Automaton (PDA) — Formal Definition

A PDA is a 6-tuple $(Q, \Sigma, \Gamma, \delta, q_0, F)$, where

- Q is a finite set called the states,
- Σ is a finite set called the input alphabet,
- Γ is a finite set called the stack alphabet,
- $\delta : Q \times \Sigma_\varepsilon \times \Gamma_\varepsilon \to \mathcal{P}(Q \times \Gamma_\varepsilon)$ is the transition function,[2]
- $q_0 \in Q$ is the starting state, and
- $F \subseteq Q$ is the set of accepting states.

The symbol ε above does not stand for the empty string (rather for a character not in $\Sigma \cup \Gamma$).
In the following we interpret the symbol ε as a character or as the empty string according to the context; specifically, ε is taken as the empty string only if the relevant alphabet does not contain the symbol ε.

[2] $X_\varepsilon := X \cup \{\varepsilon\}$.

The language accepted by a PDA

- A pushdown automaton (PDA) M accepts a string w, if there is a “computation" of M on w (see next slide) that leads to an accepting state.
- The language accepted by M, denoted $\mathcal{L}(M)$, is the set of all strings $w \in \Sigma^*$ accepted by M.
- A (non-deterministic) PDA may have many computations on a single string

Model of Computation

The following is with respect to $M = (Q, \Sigma, \Gamma, \delta, q_0, F)$.

Definition 1 ($\delta^{*,ne}$)

For $w = (w_1, \ldots, w_n) \in \Sigma_\varepsilon^n$ let $\delta^{*,ne}(w)$ be all pairs $(q, s) \in Q \times \Gamma^*$ for which exist states $r_1, \ldots, r_n \in Q$ and strings $s_0, s_1, \ldots s_n \in \Gamma^*$ s.t.:

1. $r_0 = q_0$, $r_n = q$, $s_0 = \varepsilon$ and $s_n = s$ (here ε stands for the empty string)
2. For every $i \in \{0, \ldots, n-1\}$ exist $a, b \in \Gamma_\varepsilon$ and $t \in \Gamma^*$ s.t.:
   1. $(r_{i+1}, b) \in \delta(r_i, w_{i+1}, a)$
   2. $s_i = at$ and $s_{i+1} = bt$ (here $a = \varepsilon$, or $b = \varepsilon$, is taken as the empty string).

Namely, $(q, s) \in \delta^{*,ne}(w)$ if after reading w, M can find itself in state q and stack value s.

- $\delta^*(w) = \bigcup_{w' \in (\varepsilon^* w_1 \varepsilon^* \ldots w_n \varepsilon^*)} \delta^*(w')$
- M accepts $w \in \Sigma^*$ if $\exists q \in \mathcal{F}$ such that $(q, t) \in \delta^*(w)$ for some t.

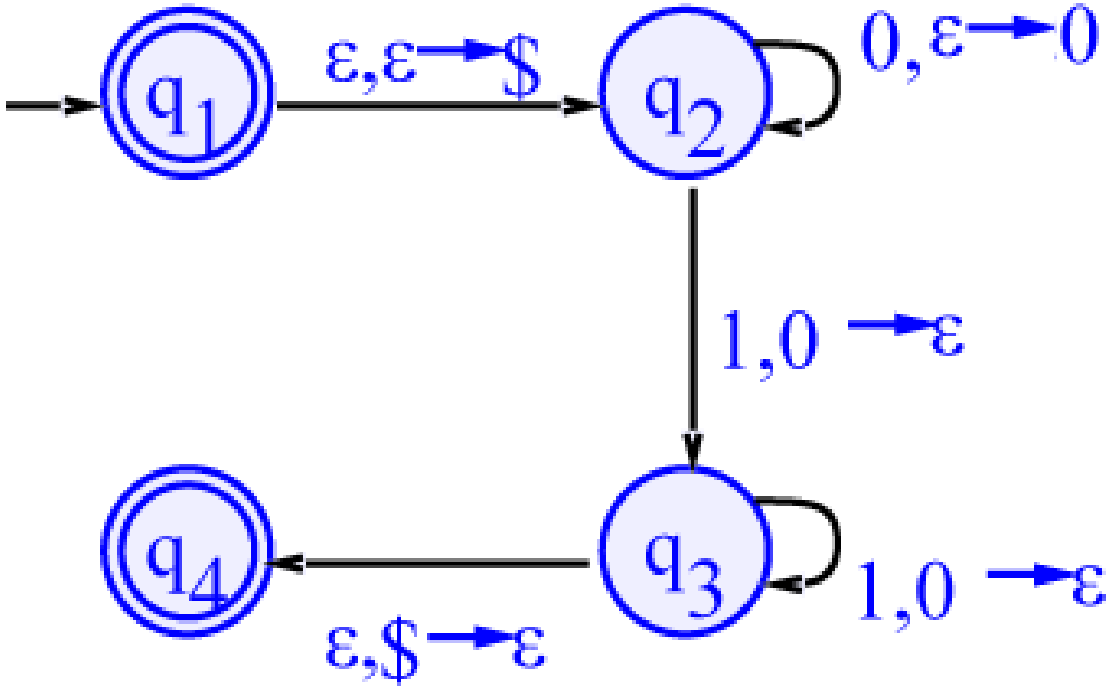
Diagram Notation

When drawing the automata diagram, we use the following notation

- Transition $a, b \to c$ from state q to q' means $(q', c) \in \delta(q, a, b)$, and informally means the automata
  - read a from input
  - pop b from stack
  - push c onto stack

- Meaning of ε transitions ((informally):
  - $a = \varepsilon$: don't read input
  - $b = \varepsilon$: don't pop any symbol
  - $c = \varepsilon$: don't push any symbol

A PDA for $\mathcal{L}_1 = \{0^n 1^n : n \geq 0\}$



Claim 2

$0011 \in L(P)$.

Proof: take

	$w'_1 = \varepsilon$	$w'_2 = 0$	$w'_3 = 0$	$w'_4 = 1$	$w'_5 = 1$	$w'_6 = \varepsilon$
$s_0 = \varepsilon$	$s_1 = \$$	$s_2 = 0\$$	$s_3 = 00\$$	$s_4 = 0\$$	$s_5 = \$$	$s_7 = \varepsilon$
$r_0 = q_1$	$r_1 = q_2$	$r_3 = q_2$	$r_4 = q_2$	$r_5 = q_3$	$r_6 = q_3$	$r_7 = q_4$

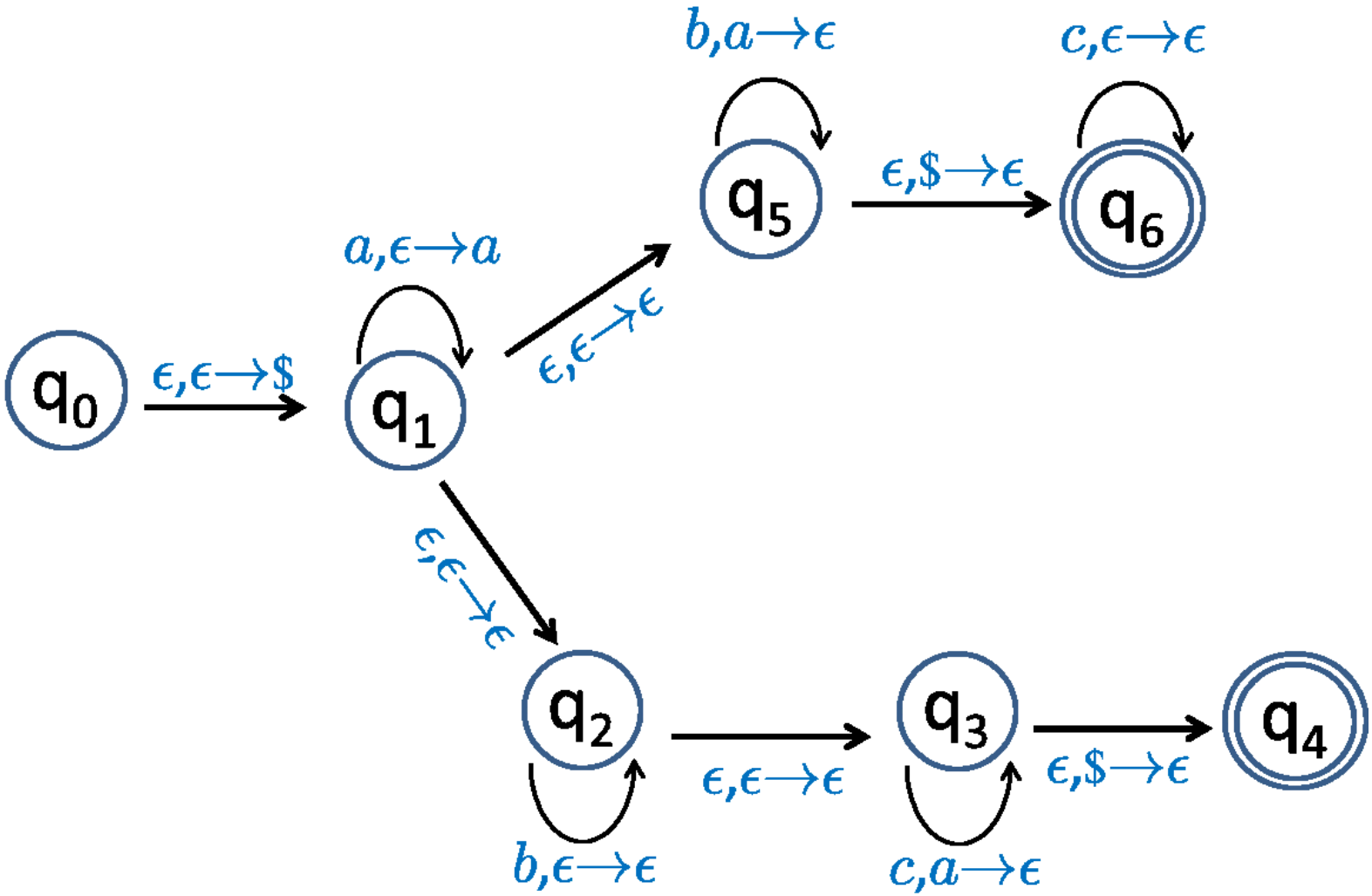
Knowing when stack is empty

It is convenient to be able to know when the stack is empty, but there is no built-in mechanism to do that.

Solution

1. Start by pushing $ onto stack.
2. When you see it again, stack is empty.

A PDA for $\mathcal{L}_2 = \{a^i b^j c^k : i = j \vee i = k\}$



A PDA for $\mathcal{L}_2 = \{a^i b^j c^k : i = j \vee i = k\}$, cont.

- Non-determinism is essential here!
- Unlike finite automata, non-determinism does add power.
- But we saw deterministic algorithm to deicide any CFL (and as we see later, CFLs are exactly the languages decided by DFAs)!
- How to prove that non-determinism adds power?

- Does not seem trivial or immediate.
- Another example: $\mathcal{L} = \{x^n y^n : n \geq 0\} \cup \{x^n y^{2n} : n \geq 0\}$ is accepted by a non-deterministic PDA, but not by a deterministic one.

PDA Languages

The Push-Down Automata Languages, $\mathcal{L}_{\text{PDA}}$, is the set of all languages that can be described by some PDA:

- $\mathcal{L}_{\text{PDA}} = \{\mathcal{L}(M) : M \text{ is a PDA}\}$

It is immediate that $\mathcal{L}_{\text{PDA}} \supsetneq \mathcal{L}_{\text{DFA}}$: every DFA is just a PDA that ignores the stack.

- $\mathcal{L}_{\text{CFG}} \subseteq \mathcal{L}_{\text{PDA}}$?
- $\mathcal{L}_{\text{PDA}} \subseteq \mathcal{L}_{\text{CFG}}$?
- $\mathcal{L}_{\text{PDA}} = \mathcal{L}_{\text{CFG}}$!!!

Part II

Equivalence Theorem

The CFG–PDA Equivalence Theorem

Theorem 3

$\mathcal{L}_{\text{PDA}} = \mathcal{L}_{\text{CFG}}$: A language is context free if and only if some pushdown automata accepts it.

This time (unlike the regular expression vs. regular languages theorem), both the proof "if" part and of the "only if" part are non trivial.

Proof sketch follows.

CFL $\Longrightarrow$ PDA

Lemma 4

$\mathcal{L}_{\text{CFG}} \subseteq \mathcal{L}_{\text{PDA}}$ *: If a language is context free, then some PDA accepts it.*

- Let $\mathcal{L}$ be a context-free language, and let $G = (V, \Sigma, R, S)$ be context-free grammar for $\mathcal{L}$
- We build a PDA $P = (Q, \Sigma, \Gamma, \delta, q_0, F)$, such that on input w it "figures out" if there is a derivation of w using G.

Question 5

How does P figure out which substitution to make?

Answer: It guesses.

Simplifying Assumptions

1. In a single move, a PDA can push a whole word (from some fixed set) into the stack (first letter at the top)

   Can we justify it?

2. When deriving a word form a CFL, we always substitute the left most variable

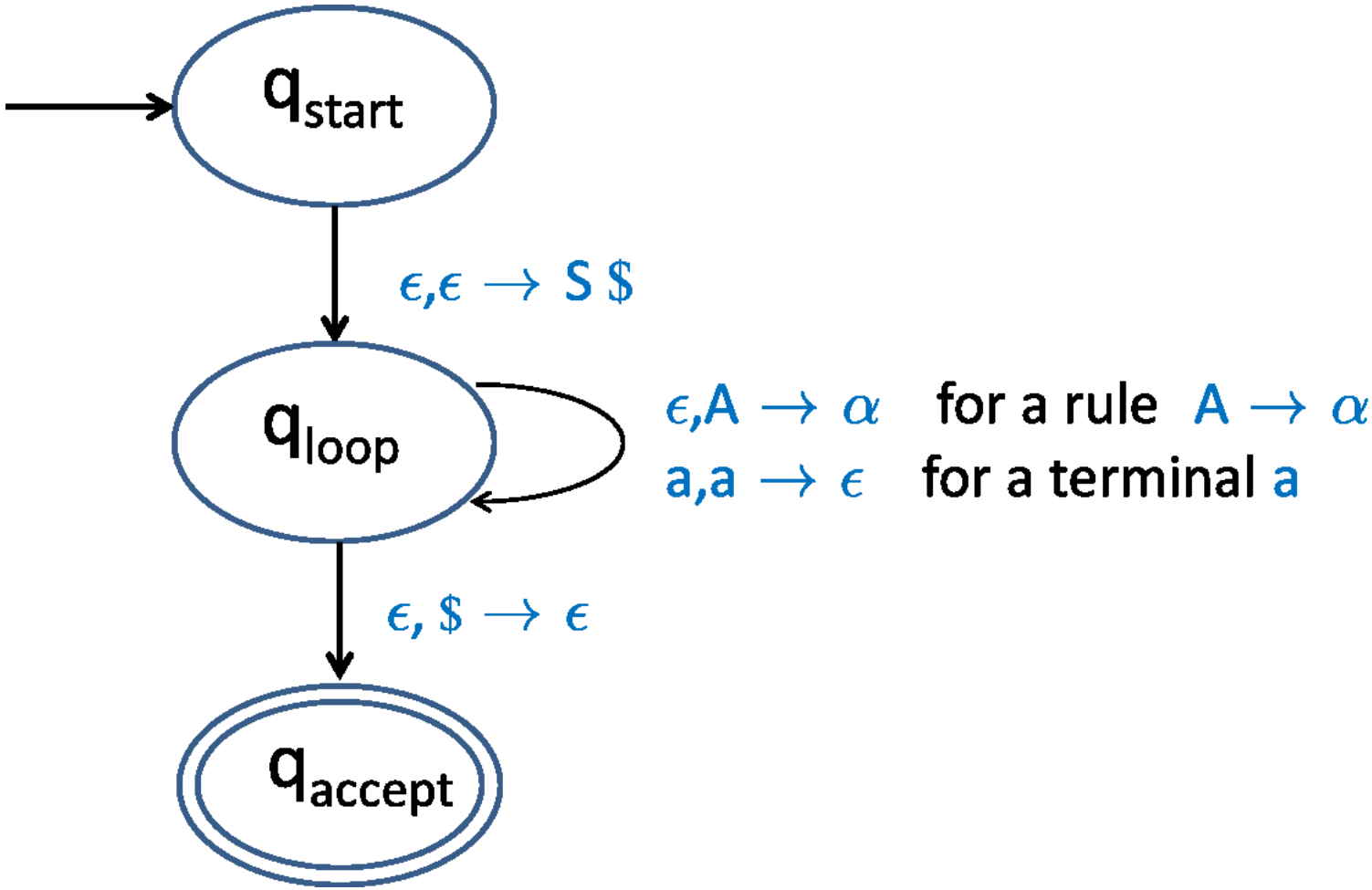
   Does it change the derived language?

Informal Description of P

Algorithm 6 (P)

1. Push $S\$$ on stack
2. While top of the stack t is not $\$$:
3. If t is variable A,
   (non-deterministically) select rule $A \to \alpha$ and substitute.
4. If t is a terminal a,
   read next input and compare; `Reject` if different.
5. `Accept` if end of input and stack is empty

State Diagram for P



Claim: $\mathcal{L}(P) = \mathcal{L}(G)$

Claim 7

$S \stackrel{*}{\rightarrow} \alpha$ iff $\alpha = \alpha_1\alpha_2$ such that $(q_{loop}, \alpha_2\$) \in \delta^*(\alpha_1)$.

Does the above yields that $\mathcal{L}(P) = \mathcal{L}(G)$?

$S \stackrel{*}{\rightarrow} \alpha \implies \alpha = \alpha_1\alpha_2$ such that $(q_{loop}, \alpha_2\$) \in \delta^*(\alpha_1)$

Proof by induction on the number of derivations steps used to yield α from S.

- 0 derivation steps: hence $\alpha = S$. Since $(q_{loop}, S\$) \in \delta^*(\varepsilon)$, the proof follows for $\alpha_1 = \varepsilon$ and $\alpha_2 = S$.
- Assume $S \stackrel{*}{\rightarrow} \alpha$ in $k > 0$ derivation steps, and let α' be the string derived by the first $(k-1)$ steps.
- By i.h $\alpha' = \alpha'_1\alpha'_2$ such that $(q_{loop}, \alpha'_2\$) \in \delta^*(\alpha'_1)$
- Write $\alpha'_2 = w_1 A w_2$ where A is the left most variable in α'_2.
- The k'th derivation step replaces this occurrence of A with a string t (why?)
- It is easy to see that $(q_{loop}, tw_2\$) \in \delta^*(\alpha'_1 w_1)$
- To complete the proof take $\alpha_1 = \alpha'_1 w_1$ and $\alpha_2 = tw_2$.

$\alpha = \alpha_1\alpha_2$ **such that** $(q_{loop}, \alpha_2\$) \in \delta^*(\alpha_1) \implies S \stackrel{*}{\to} \alpha$

Proof by induction on the number of steps used by P to process α_1.

- A single step: $\alpha_1 = \varepsilon$ and $\alpha_2 = S\$$, and the proof follows since $S \stackrel{*}{\to} S$.
- Assume α_1 was processed in $k > 1$ steps, and let α'_1 and α'_2 be the input string read and the stack value *before* the last step
- Note that $(q_{loop}, \alpha'_2\$) \in \delta^*(\alpha'_1)$.
- By i.h $S \stackrel{*}{\to} \alpha' = \alpha'_1\alpha'_2$.
- In case the k'th move of P is reading an input character, then $\alpha_1\alpha_2 = \alpha'_1\alpha'_2$, and therefore $S \stackrel{*}{\to} \alpha_1\alpha_2$
- Otherwise, $\alpha'_1 = \alpha_1$, $\alpha'_2 = Aw$ and $\alpha_2 = tw$ for some rule $A \to w \in R$
- Hence $S \stackrel{*}{\to} \alpha_1\alpha_2$

PDA $\Longrightarrow$ CFL

Lemma 8

$\mathcal{L}_{\text{PDA}} \subseteq \mathcal{L}_{\text{CFG}}$: If a PDA accepts a language then it is context free.

We prove the lemma by constructing a CFG G for a language $\mathcal{L}$ accepted by a PDA P

Let $P = (Q, \Sigma, \Gamma, \delta, q_0, F)$. We assume wlg. that:

- A single accepting state $q_a \in F$.
- P empties the stack before accepting
- Each transition either pops or pushes

Can we justify the above?

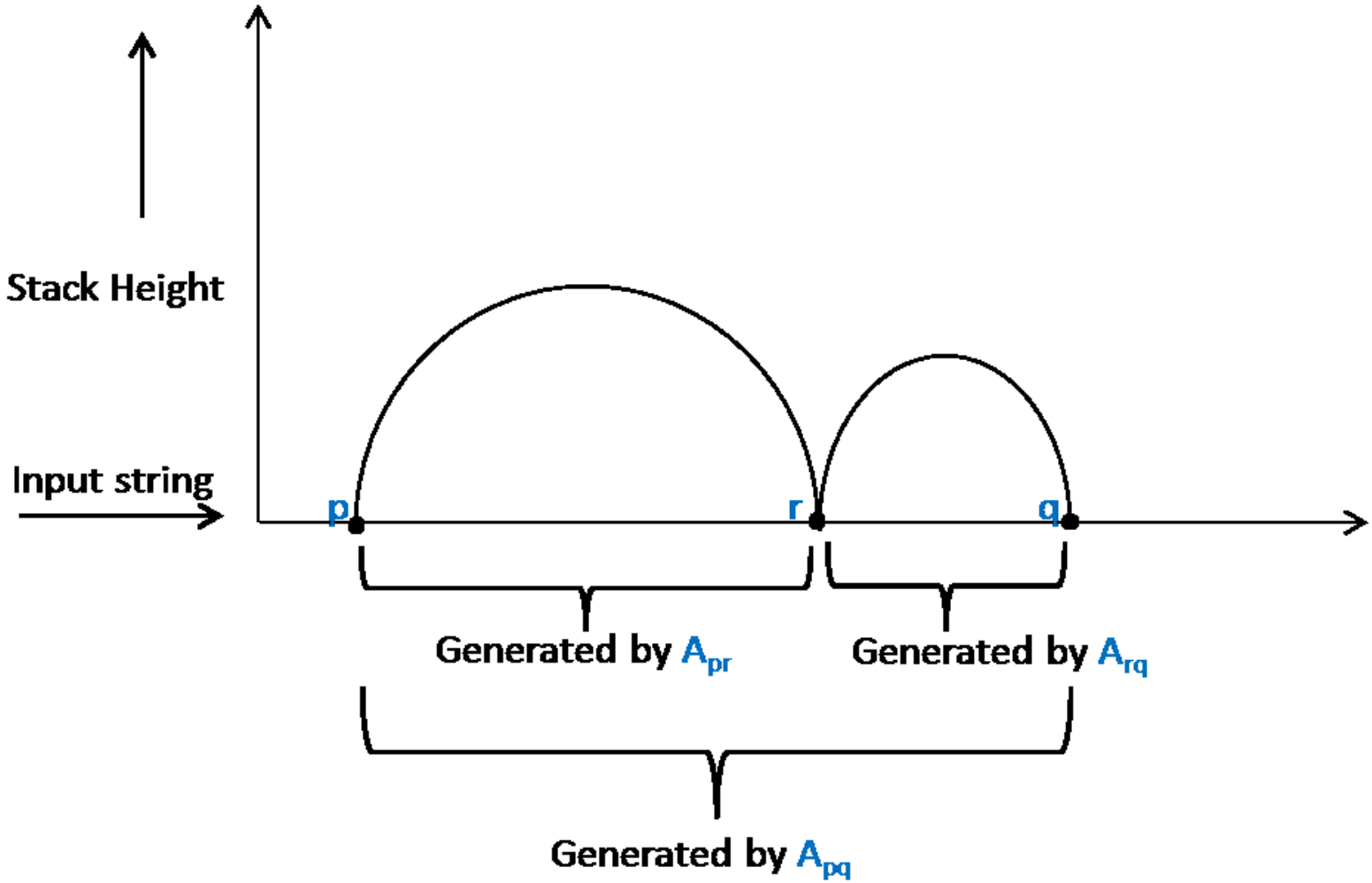
Defining $G = (V, \Sigma, R, S)$

- $V = \{A_{pq} : p, q \in Q\}$

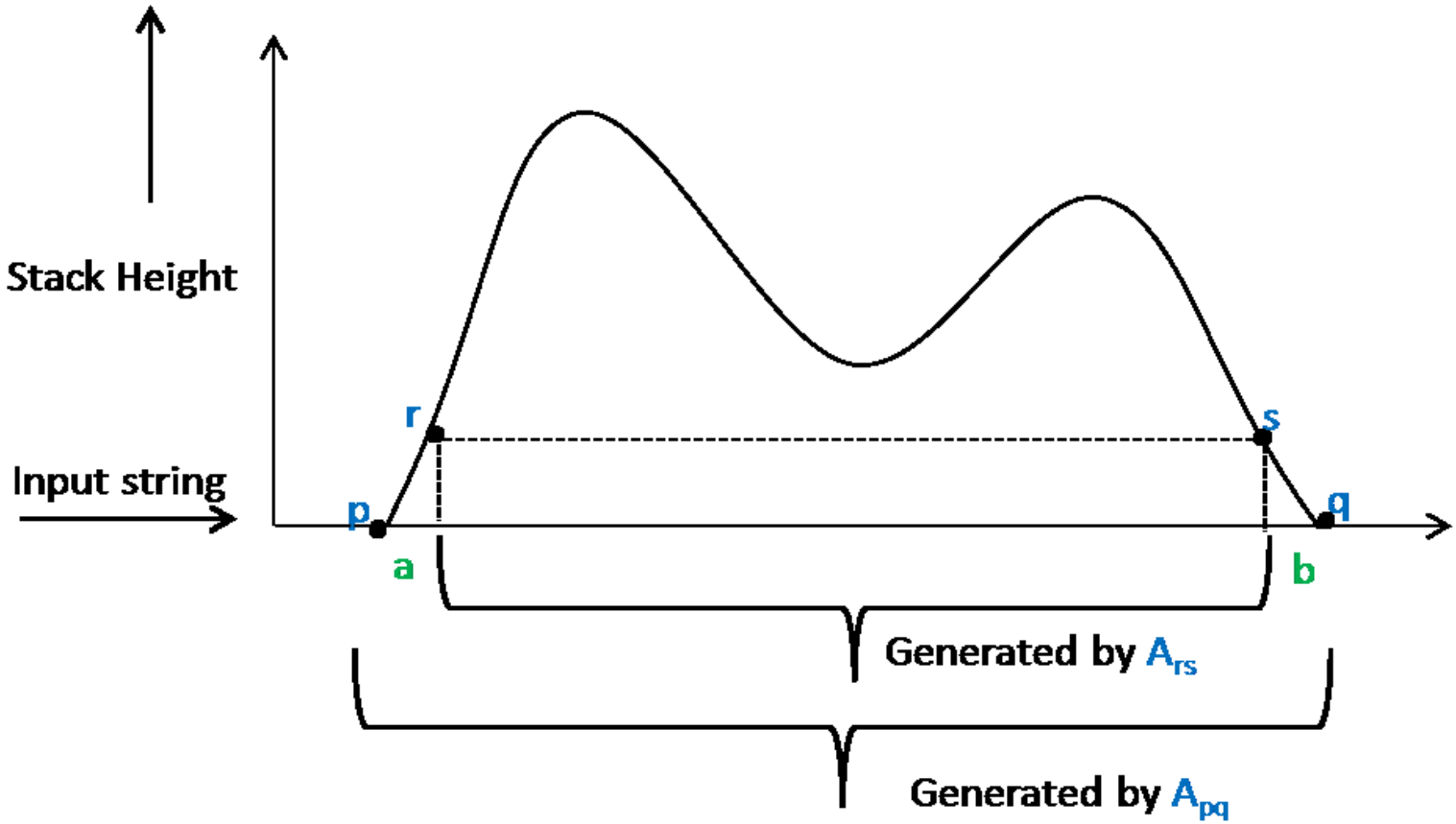
  Idea: A_{pq} will generate all strings that take P from p with an empty stack, to q with an empty stack
- $S = A_{q_0, q_a}$
- Initially $R = \emptyset$ and
  1. Add $\{A_{pq} \to A_{p,r}A_{r,q} : p, q, r \in Q\}$ to R
  2. Add $\{A_{qq} \to \varepsilon : q \in Q\}$ to R
  3. For all $p, r, s, q \in Q$, $a, b \in \Sigma_\varepsilon$ and $t \in \Gamma$ such that
     1. $(r, t) \in \delta(p, a, \varepsilon)$ and
     2. $(q, \varepsilon) \in \delta(s, b, t)$

     add $A_{pq} \to aA_{r,s}b$ to R

PDA Computation corresponding to $A_{pq} \rightarrow A_{p,r}A_{r,q}$



PDA Computation corresponding to $A_{pq} \rightarrow aA_{r,s}b$



Claim: $\mathcal{L}(G) = \mathcal{L}(P)$

For $p \in Q$, $w \in \Sigma^*$ and $t \in \Gamma_\varepsilon^*$, let $\delta^*(p, w, t)$ be the natural generalization of the single-input δ^* where the computation starts from state p and stack t (and not necessarily from q_0 and ε).

Claim 9

$A_{pq} \stackrel{*}{\rightarrow} w \in \Sigma^*$ iff $(q, \varepsilon) \in \delta^*(p, w, \varepsilon)$

Proof by induction on the number of derivation rules/ transitions

Part III

Closure Properties

Simple Closure Properties of Context-Free Languages

- CFL's are closed under
  - Union: $S \rightarrow S_1 \mid S_2$
  - Concatenation: $S \rightarrow S_1 S_2$
  - Star: $S_{new} \rightarrow \varepsilon \mid S_{old} \mid S_{new} S_{new}$
- What about complement and intersection?

Intersection

$$S_1 \to A_1 B_1 \qquad S_2 \to A_2 B_2$$
$$A_1 \to 0A_1 1 | \varepsilon \qquad A_2 \to 0A_2 | \varepsilon$$
$$B_1 \to 2B_1 | \varepsilon \qquad B_2 \to 1B_2 2 | \varepsilon$$

$$\mathcal{L}_1 = 0^n 1^n 2^* \qquad \mathcal{L}_2 = 0^* 1^n 2^n$$

- $\mathcal{L}_1 \cap \mathcal{L}_2 = 0^n 1^n 2^n$
- $\mathcal{L}_1$ and $\mathcal{L}_2$ are CFLs (why?),
- But $\mathcal{L}_1 \cap \mathcal{L}_2$ is not a CFL.
- But can't we run two PDA's in parallel, and accept iff both accept??
- What about intersection of a CFL with a regular language?

When CFL Intersects Regular Language

- Are the context free languages closed under intersection with a regular language?
- That is, if $\mathcal{L}_1$ is context free languages, and $\mathcal{L}_2$ is regular, must $\mathcal{L}_1 \cap \mathcal{L}_2$ be context free languages?
- Run PDA $\mathcal{L}_1$ and DFA $\mathcal{L}_2$ "in parallel" (just like the intersection of two regular languages).
- Formal details omitted (but you should be able to figure them out).

Applications

- Is $\mathcal{L} = \{w \in \{0,1,2\}^* : \#_0(w) = \#_1(w) = \#_2(w)\}$ context free?
  - $\mathcal{L} \cap 0^*1^*2^* = \{0^n1^n2^n : n \geq 0\}$ is not context free.
  - $0^*1^*2^*$ is regular.
  - Context free languages intersected with a regular languages are context free.
  - So $\mathcal{L}$ is not a context free language
- This could also be established using pumping lemma, but proof above is more elegant.

Complementation

The fact that CFLs are not closed under intersection, but are closed under union, implies they are not closed under complementation, as $\mathcal{L}_1 \cap \mathcal{L}_2 = \overline{\overline{\mathcal{L}_1} \cup \overline{\mathcal{L}_2}}$.

We give a simple example where $\mathcal{L}$ is not CFL but $\overline{\mathcal{L}}$ is.

- Take $\mathcal{L} = \{ww : w \in \{0,1\}^*\}$.
- $\mathcal{L}$ is not a CFL (why?)
- We prove that $\overline{\mathcal{L}}$ is a CFL

Complementation cont.

- For any $y \in \overline{\mathcal{L}}$, either
  - y's length is odd, or
  - y's length is even, 2ℓ, and $\exists i \geq 1$ such that $y_i \neq y_{\ell+i}$.
- It suffice to construct a PDA/CFG for $\overline{\mathcal{L}}_{even}$ – the even length members of $\overline{\mathcal{L}}$ (why?)
- Let $\overline{\mathcal{L}}^{\sigma}_{even} = \{\{0,1\}^k \sigma \{0,1\}^j \{0,1\}^k \overline{\sigma} \{0,1\}^j : k,j \geq 0\}$
- Note that $\overline{\mathcal{L}}_{even} = \overline{\mathcal{L}}^0_{even} \cup \overline{\mathcal{L}}^1_{even}$
- and that $\overline{\mathcal{L}}^{\sigma}_{even} = \{\{0,1\}^k \sigma \{0,1\}^k \{0,1\}^j \overline{\sigma} \{0,1\}^j : k,j \geq 0\}$
- CFG for $\overline{\mathcal{L}}^0_{even}$

  $S \rightarrow AB$

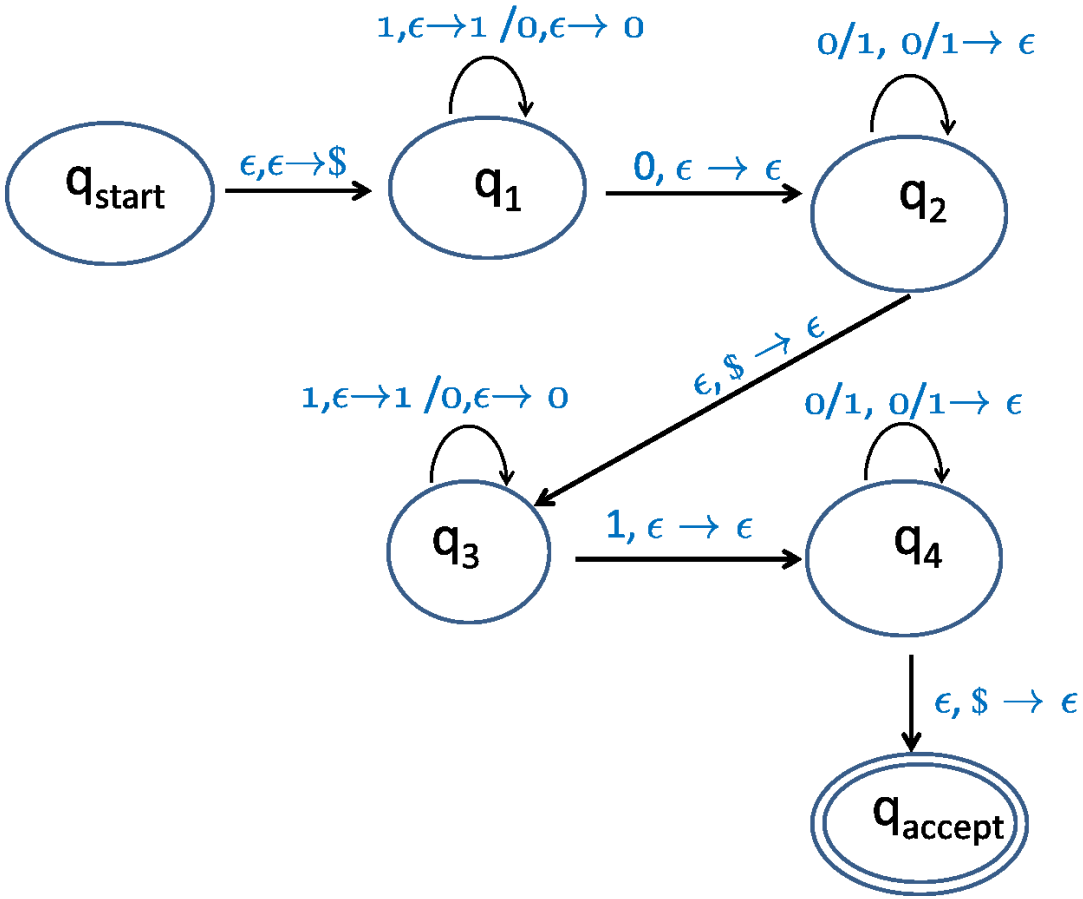
  $A \rightarrow CAC \mid 0$

  $B \rightarrow CBC \mid 1$

  $C \rightarrow 0 \mid 1$

A PDA for $\overline{\mathcal{L}}^0_{even}$

Idea: Guess $k, j \geq 0$, and accept w if it is of the form:
$\{0,1\}^k 0 \{0,1\}^k \{0,1\}^j 1 \{0,1\}^j$



Homomorphism and Inverse Homomorphism

- *Homomorphism:* replaces each letter with a word
- Example: $h(1) = aba$, $h(0) = aa$
  $h(010) = aa\,aba\,aa$
  $\mathcal{L}_1 = (01)^*$, $h(\mathcal{L}_1) = (aaaba)^*$.
- Claim: Assuming that $\mathcal{L}$ is a CFL, then so is $h(\mathcal{L})$
- Proof? use the grammar
- *Inverse homomorphism:* $h^{-1}(w) = \{x : h(x) = w\}$,
  $h^{-1}(\mathcal{L}) = \{x : h(x) \in \mathcal{L}\}$
- Example: $\mathcal{L}_2 = (ab \cup ba)^* a$, $h^{-1}(\mathcal{L}_2) = \{1\}$.
- Claim: Assuming that $\mathcal{L}$ is a CFL, then so is $h^{-1}(\mathcal{L})$
- Proof? use the automata (see next)

CFL's are Closed Under Inverse Homomorphism

- Idea: let P be a PDA for $\mathcal{L}$. On input word w, emulate $P(h(w))$
- But we cannot afford to store $h(w)$!
- Solution, compute $h(w)$ "on demand":

Algorithm 10

Input w:

1. Initialized a "buffer" Buff to $h(a)$, where a is the first letter of w
2. Emulate a running of P with Buff as its input string.

   Each time Buff is fully read by P, set Buff $= h(a)$, where a is the next letter in w (if exists)
3. `Accept` iff P does

- How do we implement Buff?

Emptiness of CFGs

Question 11

Given a CFG, G, is $\mathcal{L}(G) = \emptyset$?

In other words, is there a string generated by G?

Theorem 12

There is an algorithm that solves this problem (and always halts).

Possible approaches for a proof:

- Bad Idea: We know how to test whether $w \in \mathcal{L}(G)$ for any string w, so just try it for each w...
- Better Idea: Can the start variable generate a string of terminals?
- A more holistic approach: Can a particular variable generate a string of terminals?

Removing redundant variables and terminals of a CFG

1. Mark all terminal symbols in G.
2. Repeat until no new variable become marked:

   Mark any A where $A \to U_1 U_2 \ldots U_k$ and all U_i have already been marked.
3. Remove all unmarked variables, and any rule they appear in.
4. If S is removed, then $\mathcal{L}(G) = \emptyset$.
5. Remove any variable A not reachable from S.
6. Remove any terminal which does not appear in some rule.

Checking Emptiness

Algorithm 13

Input a CFG G

1. Remove redundant variables and terminals to get G'.
2. Return TRUE (i.e., $\mathcal{L}(G') = \emptyset$) iff S is removed

Correctness?

CFGs Fullness

Question 14

Given a CFG G, is $\mathcal{L}(G) = \Sigma^*$?

- We just saw an algorithm to determine, given a CFG G, whether $\mathcal{L}(G) = \emptyset$
- $\mathcal{L}(G) = \Sigma^*$ iff $\overline{\mathcal{L}(G)} = \emptyset$. Why not modify the algorithm so it determines emptiness of the complement?
- Unfortunately, CFGs are not closed under complement.

Fact 15

There is no algorithm to solve CFG fullness.

- We are not prepared to prove this remarkable fact (yet).

Finiteness of CFGs

Question 16

Given a CFG G, is $|\mathcal{L}(G)|$ finite?

1. Remove redundant variables and terminals.
2. Turn into a CNF form
3. Create a graph C whose nodes are variables and its directed edges are derivations.
4. Return TRUE iff C has a no cycles.

Correctness?

CFGs Inherent Ambiguity

Question 17

Given a CFG G, is $\mathcal{L}(G)$ inherently ambiguous?

This means that for any CFG that generates $\mathcal{L}(G)$, there is a word in the language with two different parse trees.

Fact 18

There is no algorithm to solve CFG inherent ambiguity.

- We will not prove this fact, yet you want to know it to put things in context.

When Are Two CFGs equivalent?

Question 19

Is there an algorithm to solve this problem?

A Short Summary

- Regular Languages $\equiv$ Finite Automata.
- Context Free Languages $\equiv$ Push Down Automata.
- Closure properties of regular languages and of CFLs.
- Most algorithmic problems for finite automata are solvable.
- Some algorithmic problems for finite automata are not solvable.
- Pumping lemmata for both classes of languages.
- There are additional languages out there.

The View Over The Horizon