Theory of Computation

Course note based on *Computability, Complexity, and Languages: Fundamentals of Theoretical Computer Science*, 2nd edition, authored by Martin Davis, Ron Sigal, and Elaine J. Weyuker.

course note prepared by

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About This Course Note

- It is prepared for the course *Theory of Computation* taught at the National Taiwan University in Spring 2010.
- It follows very closely the book *Computability, Complexity, and Languages: Fundamentals of Theoretical Computer Science*, 2nd edition, by Martin Davis, Ron Sigal, and Elaine J. Weyuker. Morgan Kaufmann Publishers. ISBN: 0-12-206382-1.
- It is available from Tyng-Ruey Chuang's web site:

  `http://www.iis.sinica.edu.tw/~trc/`

1 Regular Languages (9)

1.1 Kleene's Theorem (9.5)

Characterizations of Regular Languages

We now show that the class of regular languages can be characterized as the class of all languages obtained from finite languages using the operations $\cup, \cdot, ^*$ a finite number of times. We will see that there are other characterizations of regular languages as well.

Definitions of $L_1 \cdot L_2$ and L^*

Definition. Let $L_1, L_2 \subseteq A^*$. Then we write

$$L_1 \cdot L_2 = L_1 L_2 = \{uv \mid u \in L_1 \text{ and } v \in L_2\}.$$

□. **Definition.** Let $L \subseteq A^*$. Then we write

$$L^* = \{u_1 u_2 \dots u_n \mid n \geq 0, u_1, u_2, \dots, u_n \in L\}.$$

□. Note that, for L^*,

- $0 \in L^*$
- The notation of A^* is consistent with the definition of L^*.

$L \cdot \tilde{L}$

Theorem 5.1. If $L, \tilde{L}$ are regular languages, then $L \cdot \tilde{L}$ is regular. *Proof.* Let $\mathscr{M}$ and $\tilde{\mathscr{M}}$ be dfas that accept L and $\tilde{L}$ respectively. The two are distinct but use the same alphabet. We now construct a ndfa $\dot{\mathscr{M}}$ by "gluing together" the two dfas. We define

- the set of states $\dot{Q} = Q \cup \tilde{Q}$
- the transition function $\dot{\delta}$ by

$$\dot{\delta}(q,s) = \begin{cases} \{\delta(q,s)\} & \text{if } q \in Q - F \\ \{\delta(q,s)\} \cup \{\tilde{\delta}(\tilde{q}_1, s)\} & \text{if } q \in F \\ \{\tilde{\delta}(q,s)\} & \text{if } q \in \tilde{Q} \end{cases}$$

- the set of final states

$$\dot{F} = \begin{cases} F \cup \tilde{F} & \text{if } 0 \in \tilde{L} \\ \tilde{F} & \text{if } 0 \notin \tilde{L} \end{cases}$$

Clearly, $L \cdot \tilde{L} = L(\dot{\mathscr{M}})$, so that $L \cdot \tilde{L}$ is regular. □.

L^*

Theorem 5.2. If L is a regular languages, then so is L^*. *Proof.* Let $\mathscr{M}$ be a nonrestarting dfa that accept L. We now construct a "looping" ndfa $\tilde{\mathscr{M}}$ with the same states and initial state as $\mathscr{M}$, and accepting state q_1. The transition function $\tilde{\delta}$ is defined as follows:

$$\tilde{\delta}(q,s) = \begin{cases} \{\delta(q,s)\} & \text{if } \delta(q,s) \notin F \\ \{\delta(q,s)\} \cup \{q_1\} & \text{if } \delta(q,s) \in F \end{cases}$$

That is, whenever $\mathscr{M}$ would enter an accepting state, $\tilde{\mathscr{M}}$ will enter either the corresponding accepting state or the initial state. Clearly, $L^* = L(\tilde{\mathscr{M}})$, so that L^* is a regular language. □.

Kleene's Theorem

Theorem 5.3. A language is regular if and only if it can be obtained from finite languages by applying the three operators $\cup, \cdot, ^*$ a finite number of times. *Proof.* ($\Longleftarrow$) Every finite language is regular. The three operators build regular languages from regular languages. Therefore, by induction on the number of applications of $\cup, \cdot, ^*$, any language obtained from finite languages by applying these operators a finite number of times is regular. ($\Longrightarrow$) Let $L = L(\mathscr{M})$ where $\mathscr{M}$ is a dfa with states $q_1, \ldots, q_n$. As usual, q_1 is the initial state, F the set of accepting states, δ the transition function, and $A = \{s_1, \ldots, s_m\}$ the alphabet.

We define the sets $R^k_{i,j}$, for all $i, j > 0, k \geq 0$, as follows:

$$R^k_{i,j} = \{x \in A^* \quad | \quad \delta^*(q_i, x) = q_j \text{ and, as it moves across } x, \\ \mathscr{M} \text{ passes through no state } q_l \text{ with } l > k\}$$

Kleene's Theorem, Continued

Proof (continued). We observe that

$$\begin{aligned} R^0_{i,i} &= \{0\} \\ R^0_{i,j} &= \{a \in A \mid \delta(q_i, a) = q_j\}, \text{ for } i \neq j \end{aligned}$$

Now, to process any string of length > 1, $\mathscr{M}$ will pass through some intermediate state $q_l, l \geq 1$. We can write

$$R^{k+1}_{i,j} = R^k_{i,j} \cup (R^k_{i,k+1} \cdot (R^k_{k+1,k+1})^* \cdot R^k_{k+1,j})$$

In addition, $R^k_{i,j}$ is regular for for all i, j, k. This is proved by an induction on k. For $k = 0$, $R^0_{i,j}$ is finite hence regular. Assuming the result known for k, ($\Longleftarrow$) yields the result for $k+1$. Finally, we note that

$$L(\mathscr{M}) = \bigcup_{q_j \in F} R^n_{1,j}$$

and we conclude the proof. □

Regular Expressions

For an alphabet $A = \{ s_1, s_2, \ldots, s_k \}$, we define the corresponding alphabet

$$\mathbf{A} = \{ \mathbf{s_1}, \mathbf{s_2}, \ldots \mathbf{s_k}, \mathbf{0}, \mathbf{\emptyset}, \cup, \cdot, ^*, (,) \}.$$

The class of *regular expressions* on A is then defined to be the subset of $\mathbf{A}^*$ determined by the following:

1. $\mathbf{\emptyset}, \mathbf{0}, \mathbf{s_1}, \mathbf{s_2}, \ldots \mathbf{s_k}$ are regular expressions.
2. If α and β are regular expressions, then so is $(\alpha \cup \beta)$.

3. If α and β are regular expressions, then so is $(\alpha \cdot \beta)$.

4. If α is a regular expression, then so is $\alpha^{\boldsymbol{*}}$.

5. No expression is regular unless it can be generated using a finite number of applications of 1–4.

Semantics of Regular Expressions

For each regular expression γ, we define a corresponding regular language $\langle\gamma\rangle$ by recursion according to the following rules:

$$\begin{aligned}
\langle \mathbf{s_i} \rangle &= \{s_i\} \\
\langle \mathbf{0} \rangle &= \{0\} \\
\langle \boldsymbol{\emptyset} \rangle &= \emptyset \\
\langle (\alpha \cup \beta) \rangle &= \langle \alpha \rangle \cup \langle \beta \rangle \\
\langle (\alpha \cdot \beta) \rangle &= \langle \alpha \rangle \cdot \langle \beta \rangle \\
\langle \alpha^{\boldsymbol{*}} \rangle &= \langle \alpha \rangle^*
\end{aligned}$$

When $\langle\gamma\rangle = L$, we say that the regular expression γ *represents* L.

Regular Expressions, Examples

$$\begin{aligned}
\langle (\boldsymbol{a} \cdot (\boldsymbol{b}^* \cup \boldsymbol{c}^*)) \rangle &= \{ab^{[n]} \mid n \geq 0\} \cup \{ac^{[m]} \mid m \geq 0\} \\
\langle (\mathbf{0} \cup (\boldsymbol{a} \cdot \boldsymbol{b})^*) \rangle &= \{(ab)^{[n]} \mid n \geq 0\} \\
\langle ((\boldsymbol{c}^* \cdot \boldsymbol{b}^*)) \rangle &= \{c^{[m]}b^{[n]} \mid m, n \geq 0\}
\end{aligned}$$

Finite Subsets of A^*

Theorem 5.4. For every finite subset L of A^*, there is a regular expressions γ on A such that $\langle\gamma\rangle = L$. *Proof.* We need only to consider the following:

- If $L = \emptyset$, then $L = \langle \boldsymbol{\emptyset} \rangle$.
- If $L = 0$, then $L = \langle \mathbf{0} \rangle$.
- If $L = \{x\}$, where $x = s_{i_1} s_{i_2} \ldots s_{i_l}$, then

$$L = \langle (\boldsymbol{s_{i_1}} \cdot (\boldsymbol{s_{i_2}} \cdot (\boldsymbol{s_{i_3}} \ldots \boldsymbol{s_{i_l}}) \ldots)) \rangle.$$

- If L has more than one elements. Assuming the result is known for languages of k elements, let L have $k+1$ elements. Then we can write $L = L_1 \cup \{x\}$, where $x \in A^*$ and L_1 contains k elements. By induction hypothesis, there is a regular expression α such that $\langle \alpha \rangle = L_1$. By the above, there is regular expression β such that $\langle \beta \rangle = \{x\}$. Then we have
$$\langle (\alpha \cup \beta) \rangle = L_1 \cup \{x\} = L$$

□

Kleene's Theorem — Second Version

Theorem 5.5. A language $L \subseteq A^*$ is regular if and only if there is a regular expression γ on A such that $\langle \gamma \rangle = L$. *Proof.* ($\Longleftarrow$) For any regular expression γ, the regular language $\langle \gamma \rangle$ is built up from finite languages by applying $\cup, \cdot, ^*$ a finite number of times, so $\langle \gamma \rangle$ is regular by the Kleene's theorem. ($\Longrightarrow$) If a regular language L is finite, then by Theorem 5.4, there is a regular expression γ such that $\langle \gamma \rangle = L$. Otherwise, by Kleene's theorem, L can be obtained from certain finite languages by a finite of applications of $\cup, \cdot, ^*$.

Starting with regular expressions representing these finite languages, we then build up a regular expression representing L by simply indicating each use of the operations $\cup, \cdot, ^*$ by writing $\cup$, $\boldsymbol{\cdot}$, $\boldsymbol{^*}$, respectively, and punctuating with **(** and **)**. □

1.2 The Pumping Lemma and Its Application (9.6)

Pigeon-Hole Principle

Pigeon-Hole Principle. If $n+1$ objects are distributed among n sets, then at least one of the sets must contain at least two objects. □

Pumping Lemma

Theorem 6.1. Let $L = L(\mathscr{M})$, where $\mathscr{M}$ is a dfa with n states. Let $x \in L$, where $|x| \geq n$. Then we can write $x = uvw$, where $v \neq 0$ and $uv^{[i]}w \in L$ for all $i = 0, 1, 2, 3, \ldots$. *Proof.* Since x has at least n symbols, $\mathscr{M}$ must go through at least n state transitions. Including the initial state, this requires $\mathscr{M}$ to visit at least $n+1$ states. We conclude that $\mathscr{M}$ must visit at least one state q more than once. Then we can write $x = uvw$, where

$$\begin{aligned} \delta^*(q_1, u) &= q, \\ \delta^*(q, v) &= q, \\ \delta^*(q, w) &\in F. \end{aligned}$$

However, the loop starting and ending at q can be repeated any number of times and $\mathscr{M}$ still reachs the accepting states. It is clear that

$$\delta^*(q_1, uv^{[i]}w) = \delta^*(q_1, uvw) \in F.$$

Hence $uv^{[i]}w \in L$. □

Applications of The Pumping Lemma, I

Theorem 6.2. Let $\mathscr{M}$ be a dfa with n states. Then, if $L(\mathscr{M}) \neq \emptyset$, there is a string $x \in L(\mathscr{M})$ such that $|x| < n$. *Proof.* Let x be a string in $L(\mathscr{M})$ of the shortest possible length. Suppose $|x| \geq n$. By the pumping lemma, $x = uvw$, where $v \neq 0$ and $uw \in L(\mathscr{M})$. Since $|uw| < |x|$, this is a contradiction. Thus $|x| < n$. □ This theorem shows how to test a given dfa $\mathscr{M}$ to see whether the language it accepts is empty! We need only "run" $\mathscr{M}$ on all strings of length less than the number of states of $\mathscr{M}$. If none is accepted, we then conclude $L(\mathscr{M}) = \emptyset$.

Applications of The Pumping Lemma, II

Theorem 6.4. Let $\mathscr{M}$ be a dfa with n states. Then, $L(\mathscr{M})$ is infinite if and only if $L(\mathscr{M})$ contains a string x such that $n \leq |x| < 2n$. *Proof.* ($\Longrightarrow$) Let $L(\mathscr{M})$ be infinite. Then $L(\mathscr{M})$ must contain strings of length $\geq 2n$. Let $x \in L(\mathscr{M})$, where x has the shortest possible length $\geq 2n$. We write $x = x_1x_2$, where $|x_1| = n$ and $|x_2| \geq n$. By using the pigeon-hole principle, we can write $x_1 = uvw$, where

$$\begin{aligned} \delta^*(q_1, u) &= q, \\ \delta^*(q, v) &= q, \quad \text{with } 1 \leq |v| \leq n, \\ \delta^*(q, wx_2) &\in F. \end{aligned}$$

Thus $uwx_2 \in L(\mathscr{M})$, and $|x| > |uwx_2| \geq |x_2| \geq n$.

Applications of The Pumping Lemma, II, Continued

Proof (Theorem 6.4). Recall that we assume x is a shortest string of $L(\mathscr{M})$ with length at least $2n$. If $|x| = 2n$, then $|uwx_2| < |x| = 2n$. If $|x| > 2n$, then either uwx_2 becomes the shortest string of length at least $2n$ (which is a contradiction), or $|uwx_2| < 2n$. we conclude $n \leq |uwx_2| < 2n$.

($\Longleftarrow$) Let $x \in L(\mathscr{M})$ with $n \leq |x| < 2n$. By the pumping lemma, we can write $x = uvw$, where $v \neq 0$ and $uv^{[i]}w \in L(\mathscr{M})$ for all i. This shows that $L(\mathscr{M})$ is infinite. □ Theorem 6.4 shows how to test a given dfa $\mathscr{M}$ to see whether the language it accepts is finite! We need only run $\mathscr{M}$ on all strings x such that $n \leq |x| < 2n$, where $\mathscr{M}$ has n states. $L(\mathscr{M})$ is infinite just in case $\mathscr{M}$ accepts at least one of these strings.

Applications of The Pumping Lemma, III

The pumping lemma also provides us a technique for showing that given languages are not regular. For example, $L = \{a^{[n]}b^{[n]} \mid n > 0\}$ is not regular. Suppose it is, then $L = L(\mathscr{M})$, where $\mathscr{M}$ is a dfa and has m states. We will derive a contradiction by showing that there is a word $x \in L$, with $|x| > m$, such that there is no way of writing $x = uvw$, with $v \neq 0$, so that $\{uv^{[i]}w \mid i \geq 0\} \subseteq L$. Let $x = a^{[m]}b^{[m]}$. If we write $x = uvw$, with $v \neq 0$, then either $v = a^{[l_1]}$, or $v = a^{[l_1]}b^{[l_2]}$, or $v = b^{[l_2]}$, with $l_1, l_2 \leq m$. However, in each case, $uvvw \notin L$, contradicting the pumping lemma, so there can be no such dfa $\mathscr{M}$. We just show that L is not regular.