CS 373: Theory of Computation

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Normal Forms for Grammars

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If ϵ is in the language, we allow the rule $S \rightarrow \epsilon$. We will require that S does not appear on the right hand side of any rules.

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- How to convert *any context-free grammar* to an equivalent grammar in the Chomsky Normal Form
- We will start with a series of simplifications...

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- But a long intermediate string can lead to a short final string if there are ϵ-productions (rules of the form $A \to \epsilon$).
- Can we rewrite the grammar not to have ϵ-productions?

Eliminating ϵ-production

The Problem

Given a grammar G produce an equivalent grammar G' (i.e., $L(G) = L(G'))$ such that G' has no rules of the form $A \to \epsilon$, except possibly $S \to \epsilon$, and S does not appear on the right hand side of any rule.

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Note: If S can appear on the RHS of a rule, say $S \to SS$, then when there is the rule $S \to \epsilon$, we can again have long intermediate strings yielding short final strings.

Nullable Variables

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Fixed point algorithm: Propagate the label of nullable until there is no change.

Using nullable variables

Initial Ideas

Intuition: For every variable A in G have a variable A in G' such that $A \stackrel{*}{\Rightarrow}_{G'} w$ iff $A \stackrel{*}{\Rightarrow}_{G} w$ and $w \neq \epsilon$.

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$$\alpha_i = \begin{cases} X_i & \text{if } X_i \text{ is a non-nullable variable/terminal in } G \\ X_i \text{ or } \epsilon & \text{if } X_i \text{ is nullable in } G \end{cases}$$

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- Add rule $S' \to S$. If S nullable in G, add $S' \to \epsilon$ also.

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  - $L(G') \subseteq L(G)$: For every rule $A \to w$ in G', we have $A \stackrel{*}{\Rightarrow}_G w$ (by expanding zero or more nullable variables in w to ϵ)
  - $L(G) \subseteq L(G')$: If $\epsilon \in L(G)$, then $\epsilon \in L(G')$. If $A \stackrel{*}{\Rightarrow}_G w \in \Sigma^+$, then by induction on the number of steps in the derivation, $A \stackrel{*}{\Rightarrow}_{G'} w$. Base case: if $A \to w \in \Sigma^+$, then $A \to w$.

(Proof details skipped.)

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  - $S' \rightarrow S|\epsilon$

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  - Note: $A \rightarrow a$ is not a unit production
- Can we rewrite the grammar not to have unit-productions?

Eliminating unit-productions

Given a grammar G produce an equivalent grammar G' (i.e., $L(G) = L(G')$) such that G' has no rules of the form $A \rightarrow B$ where $B \in V'$.

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\begin{array}{ll}
I \rightarrow a \mid b \mid Ia \mid Ib & T \rightarrow F \mid T * F \\
N \rightarrow 0 \mid 1 \mid N0 \mid N1 & E \rightarrow T \mid E + T \\
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But as we shall see now, they can be (safely) eliminated

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But what if the grammar has cycles of unit productions? For example, $A \to B|a$, $B \to C|b$ and $C \to A|c$. You cannot use the "look-ahead" approach, because then you will get into an infinite loop.

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2. If $\langle A, B \rangle$ is a unit pair, then add production rules $A \to \beta_1 | \beta_2 | \cdots \beta_k$, where $B \to \beta_1 | \beta_2 | \cdots | \beta_k$ are *all the non-unit production rules of B*

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3. Remove all unit production rules.

Let G' be the grammar obtained from G using this algorithm. Then $L(G') = L(G)$

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- So a leftmost derivation of w in G can be broken up into "big-steps" each consisting of zero or more unit productions on the leftmost variable, followed by a non-unit production.

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- So a leftmost derivation of w in G can be broken up into "big-steps" each consisting of zero or more unit productions on the leftmost variable, followed by a non-unit production.
- For each such "big-step" there is a single production rule in G' that yields the same result. □

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- But a grammar may have "useless" variables which do not appear in any valid derivation
- Can we identify all the useless variables and remove them from the grammar? (Note: there may still be other redundancies in the grammar.)

Useless Symbols

Definition

A symbol $X \in V \cup \Sigma$ is *useless* in a grammar $G = (V, \Sigma, S, P)$ if there is no derivation of the form $S \stackrel{*}{\Rightarrow} \alpha X \beta \stackrel{*}{\Rightarrow} w$ where $w \in \Sigma^*$ and $\alpha, \beta \in (V \cup \Sigma)^*$.

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Removing useless symbols (and rules involving them) from a grammar does not change the language of the grammar

Revisiting Useless Symbols

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  - Type 2a: X is not "generating" (i.e., no $w \in \Sigma^*$ such that $X \stackrel{*}{\Rightarrow} w$), or
  - Type 2b: α or β contains a non-generating symbol

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     - Type 2b: for all α, β such that $S \stackrel{*}{\Rightarrow} \alpha X \beta$, α or β contains a non-generating symbol.

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        - Type 2b: for all α, β such that $S \overset{*}{\Rightarrow} \alpha X \beta$, α or β contains a non-generating symbol. Then in the new grammar all such derivations disappear (because some variable in α or β is removed).

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Only remains to show how to do the two steps in this algorithm

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Fixed point algorithm: Propagate the label (generating or reachable) until no change.

The Three Simplifications, Together

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Given a grammar G, such that $L(G) \neq \emptyset$, we can find a grammar G' such that $L(G') = L(G)$ and G' has no ϵ-productions (except possibly $S \to \epsilon$), unit productions, or useless symbols, and S does not appear in the RHS of any rule.

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Note: Applying the steps in a different order may result in a grammar not having all the desired properties.

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Furthermore, G *has no useless symbols.*

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Given $G = (V, \Sigma, S, P)$, convert to CNF

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- All remaining productions are of form $A \to X_1 X_2 \cdots X_n$ where $X_i \in V' \cup \Sigma$, $n \geq 2$ (and S does not occur in the RHS). We will put these rules in the right form by applying the following two transformations:
  1. Make the RHS consist only of variables
  2. Make the RHS be of length 2.

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Let $A \to X_1 X_2 \cdots X_n$, with X_i being either a variable or a terminal. We want rules where all the X_i are variables.

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For every $a \in \Sigma$

1. Add a new variable X_a
2. In every rule, if a occurs in the RHS, replace it by X_a
3. Add a new rule $X_a \rightarrow a$

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- How do you eliminate rules of the form $A \to B_1 B_2 \ldots B_n$ where $n > 2$?
- Replace the rule by the following set of rules

$$\begin{array}{rcl} A & \to & B_1 B_{(2,n)} \\ B_{(2,n)} & \to & B_2 B_{(3,n)} \\ B_{(3,n)} & \to & B_3 B_{(4,n)} \\ & \vdots & \\ B_{(n-1,n)} & \to & B_{n-1} B_n \end{array}$$

where $B_{(i,n)}$ are "new" variables.

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3. Reduce the RHS of rules to be of length at most two. New grammar replaces $A \rightarrow BX_aX_a$ by rules $A \rightarrow BX_{aa}$, $X_{aa} \rightarrow X_aX_a$, and $B \rightarrow X_bAAX_b$ by rules $B \rightarrow X_bX_{AAb}$, $X_{AAb} \rightarrow AX_{Ab}$, $X_{Ab} \rightarrow AX_b$