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Nguyen, Van Tang; Ogawa, Mizuhito

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Alternate Stacking Technique Revisited: Inclusion Problem of Superdeterministic Pushdown Automata

Nguyen Van Tang and Mizuhito Ogawa

This paper refines the alternate stacking technique used in Greibach-Friedman’s proof of the language inclusion problem \( L(A) \subseteq L(B) \), where \( A \) is a pushdown automaton (PDA) and \( B \) is a superdeterministic pushdown automaton (SPDA). In particular, we propose a product construction of a simulating PDA \( M \), whereas the one given by the original proof encoded everything as a stack symbol. This construction avoids the need for the “liveness” condition in the alternate stacking technique, and the correctness proof becomes simpler.

1. Introduction

Recent interest in model checking makes us recall inclusion problems. Typically, the automata theoretic explanation of model checking on finite transition systems is the decidability of the inclusion problem \( L(A) \subseteq L(B) \) among finite automata, where \( A \) and \( B \) describe a model and a specification, respectively. The standard methodology for the inclusion problem is to, (1) take the complement \( L(B)^c \), (2) take the intersection between \( L(A) \) and \( L(B)^c \), and (3) check its emptiness. This also works when \( A \) is extended to a pushdown automaton (PDA), but fails when \( B \) is extended to a pushdown automaton. To our knowledge, for decidable inclusion with a general pushdown automaton \( A \), the largest class of \( B \) that is used for alternate stacking is the class of superdeterministic pushdown automata (SPDAs), proposed by Greibach and Friedman. An SPDA is a DPDA satisfying:

1. finite delay (i.e., a bounded number of \( \epsilon \)-transitions in a row can be applied to any configuration), and
2. for two configurations sharing the same control state, transitions with the same symbol lead to configurations sharing the same control state such that the length change of stacks is the same.

In Ref. 2, the authors used the alternate stacking technique to show that the inclusion problem \( L(A) \subseteq L(B) \), where \( A \) is a PDA and \( B \) is an SPDA, is decidable. The key idea of the original proof is to construct a simulating pushdown automaton \( M \) such that \( L(A) \subseteq L(B) \) if and only if \( L(M) = \emptyset \). However, the original construction encodes everything as stack symbols (in an intricate way), and thus control states and transition rules of \( M \) could not be given in details. Furthermore, to decide the emptiness of \( M \), one has to use an auxiliary procedure to check whether a configuration of the PDA \( A \) is live (i.e., whether a configuration reaches an accepting configuration) or not. These properties of their simulating PDA \( M \) lead to a complicated proof of soundness and completeness for the decision procedure.

In this paper, we refine the alternate stacking technique used in Greibach-Friedman’s proof. Basically, there are three main steps in the proof of the decidability of the inclusion problem \( L(A) \subseteq L(B) \), where \( A \) is a PDA and \( B \) is an SPDA. First, establishing Key lemma (Lemma 3.3) to find a bounded number \( k \) that is used for alternate stacking. Second, constructing a simulating PDA \( M \) by using the alternate stacking technique (Section 3). Third, based on the construction of \( M \) in the second step, proving soundness and completeness of the construction \( L(A) \subseteq L(B) \) if \( L(M) = \emptyset \) (Section 4). Our refinement contributes to the last two steps. In particular, we give a more direct product construction of the simulating PDA \( M \), which is different from the one given by the original proof, where everything is encoded as a stack symbol. This construction avoids the need for the “liveness” condition, and the correctness proof becomes simpler.

This paper is organized as follows. In Section 2, we recall the terminology, notions, and basic definitions of superdeterministic pushdown automata. Section 3 presents our refinement on the alternate stacking technique used in Ref. 2. We show the detailed construction of simulating PDA. This section also gives a simple example to illustrate our construction technique. Section 4 provides simple proof of soundness and completeness for the decision procedure, i.e., \( L(A) \subseteq L(B) \) if \( L(M) = \emptyset \). We discuss some related works on decidable inclusion problems in Section 5. Section 6 concludes the paper.
2. Superdeterministic Pushdown Automata

2.1 Pushdown Automata

Let \( \Sigma = \{a, b, c, \ldots\} \) be a finite set of letters. The set \( \Sigma^* \) denotes all finite words over \( \Sigma \). The empty word is denoted by \( \epsilon \). A subset of \( \Sigma^* \) is called a language. Given a nonempty word \( w \in \Sigma^* \) we write \( w = a_1a_2 \cdots a_n \), where \( a_i \in \Sigma \) denotes the \( i \)-th letter of \( w \) for all \( 1 \leq i \leq n \). Let denote \( \text{head}(w) \) the first letter of \( w \), i.e., \( \text{head}(w) = a_1 \). The length \( |w| \) of \( w \) is \( n \) and \( |\epsilon| = 0 \). The notation \( |\cdot| \) also denotes the cardinality of a set, the absolute value of an integer, and the size of a pushdown automaton (see definition below).

**Definition 1.** A pushdown automaton (PDA) \( A \) over an alphabet \( \Sigma \) is a tuple \( A = (Q, \Sigma, \Gamma, Z_0, \Delta, q_0, F) \), where

1. \( Q = \{p, q, r, \ldots\} \) is a finite set of control states,
2. \( \Gamma = \{X, Y, Z, \ldots\} \) is a finite set of stack symbols such that \( Q \cap \Gamma = \emptyset \), \( Z_0 \in \Gamma \) is the initial stack symbol,
3. \( \Delta \) is a finite set of transition rules of the form \( (p, X) \xrightarrow{a, \gamma} (q, \alpha) \) where \( p, q \in Q, a \in \Sigma \cup \{\epsilon\}, X \in \Gamma \), and \( \alpha \in \Gamma^* \), and \( \epsilon \notin \Gamma \) (empty input word) is a special symbol,
4. \( q_0 \) is the initial control state,
5. and \( F \subseteq Q \) is a set of final control states.

For a rule \( (p, X) \xrightarrow{a, \gamma} (q, \alpha) \in \Delta \), we call \( (p, X) \) the mode of the rule with input \( a \); if \( a = \epsilon \), this is an \( \epsilon \)-rule. If no rule is defined for \( (p, X) \) in \( Q \times \Gamma \), \( (p, X) \) is a blocking mode. If no \( \epsilon \)-rule is defined for mode \( (p, X) \) and \( (p, X) \) is not a blocking mode, we call it a reading mode. We say that a rule \( (p, X) \xrightarrow{a} (q, \alpha) \) is a push, internal, or pop rule if \( |\alpha| = 2, 1, \) or \( 0 \), respectively. A PDA is called real-time (RPDA) if \( (p, X) \xrightarrow{a} (q, \alpha) \in \Delta \) implies that \( a \neq \epsilon \). A PDA is called deterministic (DPDA) if for every \( p \in Q, X \in \Gamma \) and \( a \in \Sigma \cup \{\epsilon\} \) we have: (1) \(|\{q, \alpha\} | (p, X) \xrightarrow{a} (q, \alpha) | \leq 1 \), and (2) if \( (p, X) \xrightarrow{a} (q, \alpha) \) and \( (p, X) \xrightarrow{a} (q', \alpha') \) then \( a = \epsilon \).

Let us denote \( St = \Gamma^* \). The set \( Q \times St \) is the set of configurations of a PDA. A pair \( (p, X) \) is a configuration with mode \( (p, X) \), written mode\((p, X) = (p, X) \). The configuration \( (q_0, Z_0) \) is called initial. For a configuration \( c = (p, y) \), the control state of \( c \) is state\((c) = p \), and the stack height of \( c \) is \(|c| = |y| \).

The transition relation between configurations is defined by: if \( (p, X) \xrightarrow{a} (q, \alpha) \), then \( (p, \beta X) \xrightarrow{\alpha} (q, \beta \alpha) \) for any \( \beta \in \Gamma^* \), and we call it one-step computation. A transition \( (p, \beta X) \xrightarrow{\alpha} (q, \beta \alpha) \) is an \( \epsilon \)-transition. If \( c_1 \xrightarrow{a_1} c_2 \) and \( c_2 \xrightarrow{a_2} c_3 \), we write \( c_1 \xrightarrow{a_1 a_2} c_3 \) and call it a computation from \( c_1 \) to \( c_3 \) on the input \( uv \). For any configuration \( c \), we write \( c \xrightarrow{\epsilon} c \), and we call it a zero-step computation, where \( \tau a = \tau a = \alpha \) for all \( a \in \Sigma \). A sequence \( c_1 \xrightarrow{a_1} c_2 \xrightarrow{a_2} \cdots \xrightarrow{a_n} c_{n+1} \) of one-step computations is an \( n \)-step computation. If we have an \( n \)-step computation \( c_1 \xrightarrow{a_1} c_2 \xrightarrow{a_2} \cdots \xrightarrow{a_n} c_{n+1} \) with \( |c_1| \leq |c_i|, 1 \leq i \leq n+1 \), we write \( c_1 \xrightarrow{(a_1 \cdots a_n)} c_{n+1} \). This is a stacking computation.

A PDA \( A \) is of delay \( d \) if, whenever there is a sequence of one-step computations: \( c_1 \xrightarrow{\epsilon} c_2 \xrightarrow{\epsilon} c_3 \cdots \xrightarrow{\epsilon} c_n \), then \( n - 1 \leq d \) (i.e., at most \( d \) \( \epsilon \)-rules in a row can be applied to any configuration). A PDA \( A \) is of finite delay if \( d \) is of delay \( d \) for some \( d \geq 0 \). It is easy to see that if a PDA is of delay 0, then it is real-time.

Languages. We consider PDAs accepting a final state and an empty stack. A language accepted from a configuration \( c \) is \( L(c) = \{w \in \Sigma^* | c \xrightarrow{w} (q, \varepsilon) \in F \} \). The language accepted by a PDA \( A \) is \( L(A) = L((q_0, Z_0)) \). The PDAs \( M_1 \) and \( M_2 \) are equivalent, denoted as \( M_1 \equiv M_2 \), if they accept the same language, i.e., \( L(M_1) = L(M_2) \). Configurations \( c_1 \) and \( c_2 \) in \( M_2 \) are equivalent, denoted as \( c_1 \equiv c_2 \), if \( L(c_1) = L(c_2) \). For a configuration \( c, c \) is accessible if \( (q_0, Z_0) \xrightarrow{w} c \) for some \( w \in \Sigma^* \). \( c \) is live if \( c \xrightarrow{w} (q, \varepsilon) \) for some \( q \in F \) and some \( w \in \Sigma^* \).

2.2 Normalized Pushdown Automata

For the purpose of our work, it is convenient to use a normal form of pushdown automata.

**Definition 2.** A pushdown automaton \( A = (Q, \Sigma, \Gamma, Z_0, \Delta, q_0, F) \) is normalized if

1. for all \( p \in Q \) and \( X \in \Gamma \), \( (p, X) \) is not a blocking mode;
2. for all \( p \in Q \) and \( \delta \) in \( \delta \) of the form \( (p, X) \xrightarrow{a} (q, \alpha) \) either satisfy \( a \in \Sigma \) or all of them satisfy \( a = \epsilon \), but not both;
3. every rule in \( \delta \) is of the form \( (p, X) \xrightarrow{a} (q, \varepsilon) \), \( (p, X) \xrightarrow{a} (q, X) \), or \( (p, X) \xrightarrow{a} (q, XY) \) where \( a \in \Sigma \cup \{\epsilon\} \).

The next lemma enables us to convert an arbitrary PDA into an equivalent normalized PDA.
Lemma 1 (Nowotka-Srba\(^3\)). For every PDA (DPDA) there is a normalized PDA (DPDA) that recognizes the same language.

2.3 Superdeterministic Pushdown Automata

Superdeterministic pushdown automata (SPDAs) were first introduced by Greibach and Friedman\(^2\). In this section, we briefly recall the standard notion and key properties of SPDAs. Readers are referred to the seminal paper\(^2\) for more details.

Definition 3. A PDA \(A = (Q, \Sigma, \Gamma, Z_0, \Delta, q, F)\) is superdeterministic if it satisfies the following conditions.

1. \(A\) is deterministic and of finite delay,
2. for all accessible configurations in reading mode \(c_1, c_2, c_1', c_2'\) and \(w \in \Sigma^*\), if both of the following are satisfied:
   - \(c_1 \xrightarrow{w} c_1'\) and \(c_2 \xrightarrow{w} c_2'\),
   - then, \(\text{state}(c_1') = \text{state}(c_2')\) and \(|c_1| - |c_1'| = |c_2| - |c_2'|\).

Remark 1. In\(^2\), Greibach and Friedman considered the blocking condition on PDAs (middle, pp.677): “Unlike Valiant, we do not allow the pda to operate with empty stack (no rules \(q,e,a,p,y\)). This avoids some complications in notation but does not affect the classes of languages involved because we allow endmarkers”. In particular, the blocking condition is not an essential restriction if we use two special symbols \# (start-maker) and $ (end-marker), where \# pushes a special stack symbol, and $ pops it.

This assumption was used to prove Key lemma (Lemma 2). More precisely, it was used to show the claim (middle, pp.684\(^2\)) that: “Hidden in many of our arguments is the following consequence of determinism and acceptance by empty store. Suppose \(L(c_1) \subseteq L(c_1')\) with \(c_1\) (but not necessarily \(c_1\)) a configuration in a deterministic pda, \(c_1 \xrightarrow{w} c_1'\), and \(c_1 \xrightarrow{w} c_2\). Then \(L(c_2) \subseteq L(c_2')\).”

Definition 4. A language \(L\) is superdeterministic if there is an SPDA \(M\) such that either \(L = L(M)\) or \(L\$ = L(M)\$\) for an end-marker $.

Note that the language \(|a^n b^n| n \geq 0\) is superdeterministic. However, due to Condition 2 in Definition 3, the language \(L = \{a^n b^m | m \geq n\}\) is not accepted by any SPDA (pp.678\(^2\)). Suppose on the contrary that there is an SPDA \(M\) accepting \(L\). While reading \(a\), \(A\) pushes a symbol, and while reading \(b\), \(A\) pops the same symbol. Thus, for instance, after reading \(a^5\) and \(a^{10}\), \(A\) will be in two configurations, \(c_1\) and \(c_2\), such that \(\text{state}(c_1) = \text{state}(c_2)\). Now concatenating \(b^{10}\), \(A\) will lead to configurations \(c_1'\) (for \(a^5b^{10}\)) and \(c_2'\) (for \(a^{10}b^{10}\)). However, \(|c_1'| - |c_1| = 0 - 5 \neq |c_2'| - |c_2| = 0 - 10\). This violates the definition of SPDAs. Moreover, as shown in\(^2\), the class of superdeterministic languages (languages accepted by SPDAs) contains the generalized parenthesis languages, which is a superclass of both parenthesis languages\(^6\) and Dyck sets.

Remark 2. It is undecidable whether a given context-free language is superdeterministic. However, it is decidable whether a given PDA \(M\) is an SPDA (pp.678\(^2\)): “It is decidable whether a dpda \(M\) is finite delay (using the decidability of emptiness and finiteness for context-free grammars and the standard construction of grammars from machines), and if \(M\) is of finite delay, an upper bound \(d\) on the delay can be computed from a description of \(M\). Knowing that \(M\) is of delay \(d\), it can be determined whether or not \(M\) is superdeterministic by examining only computations \(c \xrightarrow{w} c'\) for a symbol \(a\) with \(c\) and \(c'\) in reading mode. Since it is decidable for \(q \in Q, y \in \Gamma^*\) whether there is a \(u\) in \(\Gamma^*\) with \(q,uy\) accessible, it is decidable whether a dpda is superdeterministic. It is not known if it is decidable whether a deterministic context-free language is superdeterministic, just as it is not known whether it is decidable whether a deterministic context-free language is finite-turn or one-counter\(^7\). Standard arguments show that it is undecidable whether an arbitrary context-free language is superdeterministic”.

A PDA is called one-increasing if the stack height increases by at most one per move. As is well known, each PDA can be transformed into an equivalent one-increasing PDA.

Lemma 2 (Key Lemma 3.3\(^\dagger\)). Let \(A\) be a normalized PDA, and \(B\) be a one-increasing SPDA of delay \(d\). Let \(c_1\) be a configuration in \(A\) and \(c_1'\) be an accessible configuration in \(B\) with \(L(c_1) \subseteq L(c_1')\). Suppose we have in \(A\) a computation \(c_1 \xrightarrow{(w)c_2}\), with \(c_2\) live, and in \(B\) a computation \(c_1' \xrightarrow{w} c_2'\). Then,

(1) \(|c_1'| - |c_2'| \leq k,
(2) and if \(|c_1| = |c_2|\) then \(|c_2'| - |c_1'| \leq k,\)

where,

- \(k = (d + 1)(k_1 + 1)n(m + 1)^{2k_2} + 2d,
- \(k_1 = n + 3, k_2 = 1 + 2n^2m^2(n^2 + 4),\)
The alternate stacking technique, proposed by Valiant, involves a simulation of two PDAs $A$ and $B$ using a single stack machine $M$ whose stack contents $u_1v_1\cdots u_tv_t$ encode the stack $u_1\cdots u_t$ of $A$ and $v_1\cdots v_t$ of $B$; the machine $M$ uses $u_i$ to simulate one step of $A$ and $v_i$ for one step of $B$. In the general case, the simulating machine $M$ is not a PDA. Alternate stacking “succeeds” when the stacks can be interwoven in such a way that $M$ can be implemented as a PDA.

Valiant showed that if $A$ and $B$ are nonsingular DPDAs and $L(A) = L(B)$, then the interweaving can indeed be done so that a uniform bound can be placed on the length of segments $u_i$ and $v_i$ as long as the configurations of $A$ and $B$ are live. Then the PDA $M$ can be built so that if the stack segments exceed the bound, $M$ accepts, knowing that $L(A) \neq L(B)$. Hence $L(A) = L(B)$ iff $L(M) = \emptyset$.

### 3.1 Simulating Pushdown Automata

In this subsection, we construct a simulating PDA $M$ such that $M$ will search for possible members of $L(A) \setminus L(B)$. In principle, similar to the use of the alternate stacking technique to construct $M$. In our approach, however, the control states, stack symbols, and transition rules of $M$ are defined in the form of pairs of states, stack content, and transition rules of two PDAs, respectively.

We assume that $A = (Q, \Sigma, \Gamma, Z, \Delta, q_0^A, F_A)$ is a normalized PDA, and $B = (Q, \Sigma, \Gamma, Z, \Delta, q_0^B, F_B)$ is a normalized SPDA of delay $d$ with an assumption that $0 \notin Q_B$. Let $\$1$ and $\$2$ be fresh symbols to mark the bottom of the stack of $A$ and $B$, respectively. Let $f : \Gamma_B \cup \$1\Gamma_B \rightarrow \Gamma_B$ be a function such that $f(y) = f(\$2y) = y$ for all $y \in \Gamma_B$. Let $\tau > 0$ be an integer and let us take $2\tau$ as the segment bound for simulating the stack content of $B$. Denote $\Gamma_B^\tau = \{[y], [\$2y] | y \in \Gamma_B, 0 \leq |y| \leq 2\tau\}$. A simulating PDA $M = M(A, B, r)$ can be constructed for any choice of $r$, and the next theorem, Theorem 5, will show that if the bound $r$ is appropriately selected ($r = k + 1$, where $k$ was computed from $A$ and $B$ as in Theorem 2), then we can conclude that $L(A) \subseteq L(B)$ iff $L(M(A, B, k + 1)) = \emptyset$. Formally, the simulating PDA $M = M(A, B, r)$ is constructed as follows:

**Definition 5.** A simulating PDA of $A$ and $B$ is a tuple $M = M(A, B, r) = (Q, \Sigma, \Gamma, Z, \Delta, q_0^M, F_M)$, where:

1. $Q_M = \{p_M^A\} \cup (Q_A \times Q_B) \cup (Q_A \times \{0\}) \cup (Q_A \times Q_B \times \Gamma_B^*)$ is the set of finite states,
2. $p_M^0$ is the initial state,
3. $F_M = (F_A \times (Q_B \setminus F_B)) \cup (F_A \times \{0\})$,
4. $\Gamma_M = (\Gamma_A \cup \{\$1\}) \times \Gamma_B^r$, $Z_M = (\$1, \$2)$,
5. The transition relation $\Delta_M \subseteq Q_M \times \Gamma_M \times \Sigma \times (Q_M \times \Gamma_M^r)$ is defined as follows:

**Case I:** Simulating an internal-transition of $A$ with a transition of $B$:

1. $((p_1, p_2), (X, [vZ])) \xrightarrow{a^*} ((p_1', p_2'), (X, [vy]))$ if:
   - $\begin{cases} (p_1, X) \xrightarrow{a} (p_1', X) \in \Delta_A \\ (p_2, Z) \xrightarrow{a} (p_2', y) \in \Delta_B \\ vy \neq \varepsilon, \text{ and } |f(vy)| \leq 2r \end{cases}$
2. $((p_1, p_2), (X, [vZ])) \xrightarrow{\varepsilon} ((p_1', 0), (X, [vy]))$ if:
   - $\begin{cases} (p_1, X) \xrightarrow{\varepsilon} (p_1', X) \in \Delta_A \\ (p_2, Z) \xrightarrow{\varepsilon} (p_2', y) \in \Delta_B \\ vy = \varepsilon \text{ or } |f(vy)| = 2r + 1 \end{cases}$
3. $((p_1, p_2), (X, [vZ])) \xrightarrow{a} ((p_1', 0), (X, [vZ]))$ if:
   - $\begin{cases} (p_1, X) \xrightarrow{a} (p_1', X) \in \Delta_A \\ (p_2, Z) \text{ has no rules with input } a \end{cases}$

*1 A DPDA $M$ is nonsingular if and only if there exists $m \geq 0$ such that for any two accessible configurations $(q, w')$ and $(q', w')$ where $|w| > m$, if $L((q, w') = L((q', w'))$ then $L((q', w')) = \emptyset$.

*2 For readability, we use $(\cdot, \cdot)$ to denote a configuration of the simulating PDA $M$. 
Case II: Simulating a push-transition of $A$ with a transition of $B$.

(1) $\langle (p_1, p_2), (X, [v]) \rangle \xrightarrow{a} \langle (p'_1, 0), (X, [v]) \rangle$ if:

\[
\begin{cases}
(p_1, X) \xrightarrow{a} (p'_1, X) \in \Delta_A \\
(p_2, Z) \xrightarrow{a} (p'_2, y) \in \Delta_B \\
\text{head}(vy) = $\text{\$2}$, $|f(vy)| \leq r
\end{cases}
\]

(2) $\langle (p_1, p_2), (X, [v]) \rangle \xrightarrow{a} \langle (p'_1, 0), (X, [v]) \rangle$ if:

\[
\begin{cases}
(p_1, X) \xrightarrow{a} (p'_1, XX') \in \Delta_A \\
(p_2, Z) \xrightarrow{a} (p'_2, y) \in \Delta_B \\
\text{head}(vy) \neq $\text{\$2}$, $|f(vy)| \leq r
\end{cases}
\]

(3) $\langle (p_1, p_2), (X, [v]) \rangle \xrightarrow{a} \langle (p'_1, 0), (X', [v'']) \rangle$ if:

\[
\begin{cases}
(p_1, X) \xrightarrow{a} (p'_1, XX') \in \Delta_A \\
(p_2, Z) \xrightarrow{a} (p'_2, y) \in \Delta_B \\
r < |f(vy)| \leq 2r, \\
v_y = v'v'', |v''| = r
\end{cases}
\]

(4) $\langle (p_1, p_2), (X, [v]) \rangle \xrightarrow{a} \langle (p'_1, 0), (X', [v'']) \rangle$ if:

\[
\begin{cases}
(p_1, X) \xrightarrow{a} (p'_1, XX') \in \Delta_A \\
(p_2, Z) \xrightarrow{a} (p'_2, y) \in \Delta_B \\
r < |f(vy)| \leq 2r, \\
v_y = v'v'', |v''| = r
\end{cases}
\]
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Case III: Simulating a pop-transition of $A$ with a transition of $B$:

(1) $\langle(p_1, p_2), (X, [vZ]) \rangle \xrightarrow{a} \langle(p_1', p_2', [f(v)y]), (X', [v]) \rangle$ if:

\[
\begin{aligned}
& \{ (p_1, X) \xrightarrow{a} (p_1', X X') \in \Delta_A \\
& (p_2, Z) \xrightarrow{a} (p_2', y) \in \Delta_B \\
& |f(v'f(vy))| \leq 2r
\end{aligned}
\]

(2) $\langle(p_1, p_2), (X, [vZ]) \rangle \xrightarrow{a} \langle(p_1', p_2', [f(vy)]), (X', [v']) \rangle$ if:

\[
\begin{aligned}
& \{ (p_1, X) \xrightarrow{a} (p_1', \varepsilon) \in \Delta_A \\
& (p_2, Z) \xrightarrow{a} (p_2', y) \in \Delta_B \\
& v'vy = \varepsilon \text{ or } |f(v'f(vy))| \geq 2r + 1
\end{aligned}
\]

Case IV: When stack of $A$ is empty.

(1) $\langle(p_1, p_2), (S_1, [vZ]) \rangle \xrightarrow{a} \langle(p_1, p_2, y) \rangle$ if:

\[
\begin{aligned}
& \{ (p_2, Z) \xrightarrow{a} (p_2', y) \in \Delta_B \text{ with } a \neq \varepsilon, \\
& \text{or } (p_2, Z) \text{ is blocked}
\end{aligned}
\]

Case V: When configurations of $A$ are in the reading modes, while states of $B$ have $\varepsilon$-transitions.

(1) $\langle(p_1, p_2), (X, [vZ]) \rangle \xrightarrow{a} \langle(p_1, p_2), (X, [vy]) \rangle$ if:

\[
\begin{aligned}
& (p_1, X) \text{ is in the reading mode} \\
& (p_2, Z) \xrightarrow{a} (p_2', y) \in \Delta_B, |f(vy)| \leq 2r
\end{aligned}
\]

(2) $\langle(p_1, p_2), (X, [vZ]) \rangle \xrightarrow{a} \langle(p_1, 0), (X, [\varepsilon]) \rangle$ if:

\[
\begin{aligned}
& (p_1, X) \text{ is in the reading mode} \\
& (p_2, Z) \xrightarrow{a} (p_2', y) \in \Delta_B, |f(vy)| \geq 2r + 1
\end{aligned}
\]

The starting transition:

$\langle(p_1, p_2, y) \rangle \xrightarrow{a} \langle(q_1^n, q_2^n), (S_1, [S_2]) \rangle (\Delta_A, [S_2 B])$

Before defining configurations of $M$, let us briefly explain the intuition behind its transition rules.

- Rules I (2), II (2), and V (2) are called stacking-fail transitions. Taking a stacking-fail transition, $M$ changes its control to states in the set $Q_A \times \{0\}$. After entering this set $Q_A \times \{0\}$ of states, $M$ continues simulating transitions of $A$ only by using rules I (4), II (6), or III (4).
- Rules I (3), II (5), and III (3) are used when $B$ is blocked where reading an input. In this cases, $M$ changes its control state to the set $Q_A \times \{0\}$. After entering a state in $Q_A \times \{0\}$, $M$ only simulates transitions of $A$ by using I (4), II (6), or III (4).
- Rules II (1), II (2), II (3), and II (4) are used to simulate a push-transition of $A$ with a transition of $B$, which has the same label.
- Rules III (1), III (2) are used to simulate a pop-transition of $A$.
- Rules IV (1) and IV (2) are used when the stack of $A$ is empty; in this case, $M$ simulates $\varepsilon$-transitions of $B$ using a zero-step computation of $A$. Recall that $B$ is finite delay of $d$, and thus rules IV (1) and IV (2) can be applied at most $d$ times in a sequence.
Definition 6 (Configuration). A configuration of $M$ is of the form $c = (s, (s_1, \ldots, s_{|s|}), \ldots, (t, v_1))$, where $s \in Q_M$ and $(X_1, v_1) \in \Gamma_M$ for $1 \leq i \leq t$.

Remark 3. There are three main steps in the proof of the decidability of the inclusion problem $L(A) \subseteq L(B)$, where $A$ is a PDA and $B$ is an SPDA. First, establish the Key lemma to find a bounded number $k$ that is used for alternate stacking. Second, construct a simulating PDA $M$ by using the alternate stacking technique. Third, based on the construction of $M$ in the second step, prove the soundness and completeness of the construction $L(A) \subseteq L(B)$ iff $L(M) = \emptyset$.

3.2 An Example

This subsection provides an example to illustrate our construction of the simulating pushdown automata. In the following figures, for simplicity, we describe control states of each PDA as nodes of a graph. We adopt the following conventions to represent edges: for a transition rule $(p, X) \xrightarrow{a, b, c} (q, y)$, we label the edge from $p$ to $q$ as $a, X \rightarrow y$.

Example 1. Consider two PDAs $C$ (in Fig. 1) and $D$ (in Fig. 2) over the input alphabet $\Sigma = \{a, b, c\}$, where $D$ is an SPDA. The PDA $C = \{q_0, q_1, q_2, \Sigma, \{Z, Z_C\}, Z_C, \Delta_C, \{s_0, s_2\}\}$, where $\Delta_C$ is defined as:

- $(s_0, Z_C) \xrightarrow{\epsilon} (s_0, Z_C)$
- $(s_0, Z) \xrightarrow{a} (s_0, ZZ)$
- $(s_1, \epsilon) \xrightarrow{q_1}$
- $(s_1, Z) \xrightarrow{b} (s_1, \epsilon)$
- $(s_1, Z_C) \xrightarrow{c} (s_1, Z_C)$
- $(s_1, Z_C) \xrightarrow{\epsilon} (s_2, \epsilon)$

$L(C) = \{a^n b^n c^m | n \geq 1, m \geq 1\}$.

The SPDA $D = \{q_0, q_1, q_2, q_3, \Sigma, \{Z', Z_D\}, Z_D, \Delta_D, \{q_0, q_3\}\}$, where $\Delta_D$ is defined as:

- $(q_0, Z_D) \xrightarrow{a} (q_0, Z_D)$
- $(q_0, Z_D) \xrightarrow{b} (q_1, Z_D Z')$
- $(q_1, Z') \xrightarrow{\epsilon} (q_1, Z' Z')$
In this case, \( r = 1 \) is sound and complete, i.e., Lemma 3.

**Proof.**

\[ w \overset{w}{\rightarrow} \varepsilon \]

The simulating pushdown automaton \( M = M(C, D, 1) \) is illustrated in Fig. 3. In this case, \( r = 1 \) is sufficient to consider stack symbols of the forms \( \{X, [v]\} \) with \(|v| \leq 2\). The language of \( M \) is \( L(M(C, D, 1)) \equiv \{a^m b^n c^m \mid n \geq 1, m \geq 1\} \).

4. Soundness and Completeness

In this section, we show that the construction presented in the preceding section is sound and complete, i.e., \( L(A) \subseteq L(B) \) if and only if \( L(M(A, B, k + 1)) = \emptyset \), where \( k \) was computed from \( A \) and \( B \) as in Lemma 2.

4.1 Soundness

**Lemma 3.** \( L(A) \not\subseteq L(B) \) implies \( L(M(A, B, r)) \neq \emptyset \) for all \( r \geq 1 \).

**Proof.** Let \( w \in L(A) \setminus L(B) \). It is sufficient to show that \( w \notin L(M) \). Denote \( c_{in} \) as the initial configuration of \( M \). Recall that \( A \) is normalized (by Lemma 1), there is a computation of \( A \) on every word. By the definition of transitions of \( M \), there is a computation of \( M \) on \( w \). There are three cases:

- **Case 1:** There are no computations of \( B \) on \( w \), or there is a computation of \( B \) on \( w \) but after reading \( w \), the stack of \( B \) is nonempty. By transitions of \( M \), we have \( c_{in} \overset{w}{\rightarrow} (p, 0, \varepsilon) \). Since \( w \in L(A) \), \( (p, 0) \in F_A \times \{0\} \). Thus, \( w \in L(M) \).
- **Case 2:** There is a computation of \( B \) on \( w \) leading to a configuration \( (q, \varepsilon) \), where \( q \notin F_B \). Because \( A \) accepts \( w \), there is a computation of \( M \) on \( w \) leading to a configuration \( (q, \varepsilon) \). Thus, \( w \in L(M) \).
- **Case 3:** Where simulating \( w \), the stacking fails. In this case, we have \( c_{in} \overset{w}{\rightarrow} (p, \varepsilon), p \in F_A \).

4.2 Completeness

**Lemma 4.** Let \( k \) be the number computed in Lemma 2. \( L(M(A, B, k + 1)) \neq \emptyset \) implies \( L(A) \not\subseteq L(B) \).

**Proof.** Let \( w \in L(M(A, B, k + 1)) \). Thus, there is an accepting computation of \( M(A, B, k + 1) \) on \( w \). We consider two cases of accepting configurations of \( M \).

1) **Case 1:** \( c_{in} \overset{w}{\rightarrow} (p, \varepsilon) \) where \( (p, \varepsilon) \in F_A \times (Q_B \setminus F_B) \). In this case, there is a computation of \( B \) on \( w \) leading to the configuration \( (q, \varepsilon) \). Because \( q \notin F_B \), we obtain \( w \notin L(B) \). On the other hand, on reading \( w \), \( A \) leads to the accepting configuration \( (p, \varepsilon) \), i.e., \( w \in L(A) \). Thus, \( w \in L(A) \setminus L(B) \).

2) **Case 2:** \( c_{in} \overset{w}{\rightarrow} (p, \varepsilon) \) where \( p \in F_A \). Consider two subcases. First, if \( B \) is blocked at some point on reading \( w \). In this case, there is not a computation of \( B \) on \( w \), i.e., \( w \notin L(B) \). On the other hand, on reading \( w \), \( A \) leads to the configuration \( (p, \varepsilon) \), \( p \in F_A \). Hence \( w \in L(A) \). Since \( B \) is deterministic, we have \( w \in L(A) \setminus L(B) \). The proof is completed. Second, if stacking fails at some point on simulating \( w \). In this case, to prove \( L(A) \not\subseteq L(B) \), we assume on the contrary that \( L(A) \subseteq L(B) \). We will show a contradiction. Since the stacking fails on reading \( w \), we suppose that \( w = w_1 w_2 \) such that, after reading \( w_1 \) the first time, stacking fail occurs and \( M \) is in the control \( (p_1, 0) \) with the stack content \( \{s_1, s_2\} \cdot \ldots \cdot \{s_{t-1}, s_{t-1}\} \cdot \{s_t, s_t\} \).

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Whereas after reading \( w_1 \), \( A \) is in the configuration \( c_2 = (p_1, X_1 ..., x_k) \) (\( c_2 \) is live) and \( B \) is in the configuration \( c'_2 = (p_2', f(v_1 ..., v_{i-1})) \). There are two subcases which lead to the stacking failure: either \( [v_i] = [\epsilon] \) or \( [f(v_i)] \geq 2r + 1 \).

- **If** \( [v_i] = [\epsilon] \): we have \( t \geq 2 \) and \( f(v_1 ..., v_i) \neq \epsilon \) (because, if \( t = 1 \) then \( [v_i] \) must be \( [S_2] \), and if \( f(v_1 ..., v_i) = \epsilon \) then the stack of \( B \) is empty and \( [v_i] = [S_2] \)). Since \( f(v_1 ..., v_i) \neq \epsilon \) there is at least one \( f(v_i) \neq \epsilon \). Select the “nearest” \( v_j \) such that \( f(v_j) \neq \epsilon \) and \( f(v_i) = \epsilon \) for \( j + 1 \leq i \leq t \). Consider the time when the level \( j + 1 \) of the stack is opened. Since \( f(v_j) \neq \epsilon \), this means that the rule \( \Pi \) (3) or (4) was used, and the “new” top segment at that time was \( v_{j+1}' \) with \( |v_{j+1}'| = r \). Since that time, \( M \) has not read below level \( j + 1 \). Thus, we have \( w_1 = w'w'' \), and after reading \( w' \), \( M \) is in the configuration \((p_1', p_2', (S_1, [S_2]) \cdots (X_{j-1}, [v_i]))(X_{j+1}', [v_{j+1}'])\)) encoding the configurations

\[
c_1 = (p_1, X_1 ..., X_{j-1}) \text{ of } A \text{ and } c'_1 = (p_2', f(v_1 ..., v_{j-1})'') \text{ of } B \text{ such that:}
\]

\[
c_0 \xrightarrow{w'} c_1 \text{ and } c_0' \xrightarrow{w''} c'_1 \text{ (}c_0 \text{ and } c_0' \text{ are the initial configurations of } A \text{ and } B,\text{ respectively).}
\]

Because \( L(A) \subseteq L(B) \) (by assumption) and \( B \) is deterministic, \( L(c_1) \subseteq L(c'_1) \). On the other hand, we have \( c_1 \uparrow (w'w'')c_2 \) and \( c_1' \xrightarrow{w''} c_2' \). Note that these conditions satisfy assumptions of the Key lemma (Lemma 2). However, we have \( |c'_1| - |c_1| = |v_{j+1}'| \geq r = k + 1 > k \). This contradicts Lemma 2. Hence, the assumption \( L(A) \subseteq L(B) \) is wrong. Thus, \( L(A) \not\subseteq L(B) \).

- **If** \( |f(v_i)| \geq 2r + 1 \): Consider the time when the level \( t - 1 \) of the stack is opened. At that point, one of rules \( \Pi \) (1), \( \Pi \) (2), \( \Pi \) (3), or \( \Pi \) (4) was used and the “new” top segment was \( v_{t-1}' \) with \( |v_{t-1}'| \leq r + 1 \). Since that time, \( M \) has not read below level \( t - 1 \). Thus, we have \( w_1 = w'w'' \), and after reading \( w' \), \( M \) is in the configuration \((p_1, p_2', (S_1, [S_2]) \cdots (X_{t-1}, [v_i]))(X_t', [v_{t-1}'])\)) encoding the configurations

\[
c_1 = (p_1, X_1 ..., X_{t-2}) \text{ of } A \text{ and } c'_1 = (p_2', f(v_1 ..., v_{t-2})'') \text{ of } B \text{ such that:}
\]

\[
c_0 \xrightarrow{w'} c_1 \text{ and } c_0' \xrightarrow{w''} c'_1 \text{. Because } L(A) \subseteq L(B) \text{ (by assumption) and } B \text{ is deterministic, } L(c_1) \subseteq L(c'_1) \text{. In addition, we have } c_1 \uparrow (w'w'')c_2 \text{ and } c'_1 \xrightarrow{w''} c'_2 \text{.}
\]

Note that these conditions satisfy assumptions of Lemma 2. Now, we can compute:

\[
\begin{align*}
|c_1| - |c_2| &= |X_{t-1}| - |X_{t-1}'| = 1 - 0 \\
|c'_1| - |c'_2| &= |v_{t-1}'| - |v_{t-1}'| 
\end{align*}
\]

This contradicts Lemma 2. Hence \( L(A) \not\subseteq L(B) \).

In both cases, we have, if \( L(M(A, B, k + 1)) \neq \emptyset \), then \( L(A) \not\subseteq L(B) \). The lemma is proved.

From Lemmas 3 and 4, we obtain:

**Lemma 5.** \( L(A) \subseteq L(B) \) if and only if \( L(M(A, B, k + 1)) = \emptyset \).

### 4.3 The Inclusion Problem

Let \( A = (Q, \Sigma, \Gamma, \Delta, q_0, F) \) be a PDA. The size \( |A| \) of a PDA \( A \) is defined as \( |Q| + \|\Sigma\| + \|\Gamma\| + \{\|pXq\| \mid (p, X) \xrightarrow{\epsilon} (q, \alpha) \in \Delta\} \). We obtain the same complexity class as that of the original construction.

**Theorem 1.** The inclusion problem \( L(A) \subseteq L(B) \), where \( A \) is a PDA and \( B \) is an SPD\( A \), is decidable. Furthermore, the decision procedure has time complexity bounded by \( 2^{2n(h)} \), where \( p(h) \) is a polynomial time in the size of both automata, \( h = |A| + |B| \).

**Proof.** The decidability follows from Lemma 5. We now approximate the size of \( M \). Recall that the emptiness problem can be decided in \( O(n^2) \) for any PDA of size \( n \). The stack of \( M \) is bounded by \( |\Gamma_A| \cdot |\Gamma_B|^{2k+2} \), where \( k \) is the number given in Lemma 2. The maximum number of control states of \( M \) is \( |Q_A| \cdot |Q_B| \cdot |\Gamma_B|^{2k+1} \).

The number of transitions of \( M \) is bounded by \( |Q_A|^2 \cdot |Q_B|^2 \cdot |\Sigma|^3 \cdot |\Gamma_A|^3 \cdot |\Gamma_B|^{6k+6} \). Recall that \( s = |Q_A| + |Q_B|, g = |\Gamma_A| + |\Gamma_B| \). The size of \( M \) is bounded by \( |M| \leq s^4g^{6k+6} \). Lemma 2 expresses that \( k = (d + 1)(s + 4)g^{(2^{(1+2s^2g^2(s^2+b^2))} + 2d)} \), where \( d \) is the delay of \( B \). Define \( h = s + g \), and we see that \( k \leq h^{c \cdot h^{k \cdot h^{k}}} \), for some constants \( c_1 \) and \( c_2 \). Thus, for some constant \( c_3 \), the size of \( M \) is bounded by \( h^{c_3 \cdot h^{k \cdot h^{k}}} \). Thus, the time complexity of the construction is bounded by \( 2^{2n(h)} \) for a polynomial \( p(h) \).

### 5. Related Work

SPD\( As \) were proposed by S. Greibach and E. Friedman in Refs. 2) and 4). It is shown that the acceptance condition of SPD\( As \) does strictly affect decision problems. More precisely, for SPD\( As \) accepting by final control state, the inclusion
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If we consider SPDAs accepting by a final state and an empty stack, it is shown that the language inclusion problem \( L(A) \subseteq L(B) \) is decidable for \( A \) is an arbitrary PDA, and \( B \) is an SPDA\(^2\). As far as we know, the class of SPDAs is the largest class which enjoys decidability for this inclusion problem. The main results of the inclusion problem \( L(A) \subseteq L(B) \), in which \( A \) is an arbitrary PDA and \( B \) is an SPDA, can be summarized as follows:

- This inclusion problem is undecidable if \( B \) accepting by the final state\(^4\).
- This inclusion problem is decidable if \( B \) accepting by the final state and the empty stack\(^2\).

Some works related to the inclusion problem of context-free languages have been published recently by Minamide and Tozawa\(^1,5\). In Ref. 1), Minamide and Tozawa developed two algorithms for deciding the inclusion \( L(G_1) \subseteq L(G_2) \) where \( G_1 \) is a context-free grammar and \( G_2 \) is either an XML-grammar or a regular hedge grammar. Tozawa and Minamide\(^3\) proved further that these algorithms for XML-grammars and regular hedge grammars are PTIME and 2EXPTIME, respectively. These algorithms were incorporated into the PHP string analyzer and validated several publicly available PHP programs against XHTML DTD. The languages of XML-grammars or regular hedge grammars are subclasses of generalized parenthesis languages. On the other hand, the class of languages of SPDAs contains the class of generalized parenthesis languages\(^2\). Thus, SPDAs are more expressive than XML-grammars and regular hedge grammars.

6. Conclusion

This paper refined the alternate stacking technique used in Greibach-Friedman’s proof of the language inclusion problem \( L(A) \subseteq L(B) \), where \( A \) is a PDA and \( B \) is an SPDA. The original construction encodes everything as stack symbols (in an intricate way), whereas our refinement gives a more direct product construction, and clarifies how alternate stacking works. For our construction, a proof of “liveness” is not needed, and the whole correctness proof for the decision procedure became simpler. As mentioned, the key lemma (Lemma 2) plays a crucial role in the decidable inclusion for SPDAs. However, the original proof of the key lemma\(^2\) is indeed intricate. It would be interesting to improve the proof of this lemma.

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Nguyen Van Tang received his M.S. in 2005 from Hanoi University of Science, Hanoi, Vietnam. His research interests include formal languages, real-time systems, and verification methodology, such as model checking and theorem proving.

Mizuhito Ogawa received his M.S. in 1985 and Ph.D degree in 2002, both from the University of Tokyo. He worked in NTT Basic Research Laboratories from 1985 until 2001, and in JST from 2002 to 2003. From 2003, he has been working at the Japan Advanced Institute of Science and Technology. His research interests include theory in rewriting, formal languages, combinatorics, and program verification methodology, such as theorem provers and model checkers.