On Winning Conditions of High Borel Complexity in Pushdown Games

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Abstract. In a recent paper [19, 20] Serre has presented some decidable winning conditions \( \Omega_{A_1 \triangleright \ldots \triangleright A_n \triangleright A_{n+1}} \) of arbitrarily high finite Borel complexity for games on finite graphs or on pushdown graphs.

We answer in this paper several questions which were raised by Serre in [19, 20]. We study classes \( C_n(A) \), defined in [20], and show that these classes are included in the class of non-ambiguous context free \( \omega \)-languages. Moreover from the study of a larger class \( C_\lambda(A) \) we infer that the complements of languages in \( C_n(A) \) are also non-ambiguous context free \( \omega \)-languages. We conclude the study of classes \( C_n(A) \) by showing that they are neither closed under union nor under intersection.

We prove also that there exists pushdown games, equipped with winning conditions in the form \( \Omega_{A_1 \triangleright A_2} \), where the winning sets are not deterministic context free languages, giving examples of winning sets which are non-deterministic non-ambiguous context free languages, inherently ambiguous context free languages, or even non context free languages.

Keywords: Pushdown automata; infinite two-player games; pushdown games; winning conditions; Borel complexity; context free \( \omega \)-languages; closure under boolean operations; set of winning positions.

1. Introduction

Two-player infinite games have been much studied in set theory and in particular in Descriptive Set Theory. Martin’s Theorem states that every Gale Stewart game \( G(A) \), where \( A \) is a Borel set, is determined, i.e. that one of the two players has a winning strategy [14].
In Computer Science, the conditions of a Gale Stewart game may be seen as a specification of a reactive system, where the two players are respectively a non terminating reactive program and the “environment”. Then the problem of the synthesis of winning strategies is of great practical interest for the problem of program synthesis in reactive systems. Büchi-Landweber Theorem states that in a Gale Stewart game $G(A)$, where $A$ is a regular $\omega$-language, one can decide who is the winner and compute a winning strategy given by a finite state transducer.

In [23, 16] Thomas asked for an extension of this result to games played on pushdown graphs. Walukiewicz firstly showed in [25] that one can effectively construct winning strategies in parity games played on pushdown graphs and that these strategies can be computed by pushdown transducers. Several authors have then studied pushdown games equipped with other decidable winning conditions, [4, 5, 17, 11]. Cachat, Duparc and Thomas have presented the first decidable winning condition at the $\Sigma_3$ level of the Borel hierarchy [6]. Bouquet, Serre and Walukiewicz have studied winning conditions which are boolean combinations of a Büchi condition and of the unboundedness condition which requires the stack to be unbounded, [3]. Recently Serre has given a family of decidable winning conditions of arbitrarily high finite Borel rank [19, 20]. A game between two players Adam and Eve on a pushdown graph, is equipped with a winning condition in the form $\Omega_{A_1 \triangleright \ldots \triangleright A_n \triangleright A_{n+1}}$, where $A_1, \ldots, A_n$ are deterministic pushdown automata, the stack alphabet of $A_i$ being the input alphabet of $A_{i+1}$, and $A_{n+1}$ is a deterministic pushdown automaton with a Büchi or a parity acceptance condition. Then an infinite play is won by Eve iff during this play the stack is strictly unbounded, that is converges to an infinite word $x$ and its limit $x \in L(A_1 \triangleright \ldots \triangleright A_n \triangleright A_{n+1})$, where $L(A_1 \triangleright \ldots \triangleright A_n \triangleright A_{n+1})$ is an $\omega$-language defined as follows. A word $\alpha_0$ is in $L(A_1 \triangleright \ldots \triangleright A_n \triangleright A_{n+1})$ iff: for all $1 \leq i \leq n$, when $A_i$ reads $\alpha_{i-1}$ its stack is strictly unbounded and the limit of the stack contents is an $\omega$-word $\alpha_i$; and $A_{n+1}$ accepts $\alpha_n$. Serre proved that for these winning conditions one can decide the winner in a pushdown game and that the winning strategies are effective.

We solve in this paper several questions which are raised in [19, 20]. We first study the classes $\mathbb{C}_n(A)$ which contain languages in the form $L(A_1 \triangleright \ldots \triangleright A_n \triangleright A_{n+1})$, where $A$ is the input alphabet of $A_1$. We show that these classes are included in the class of non-ambiguous context free $\omega$-languages. Moreover from the study of a larger class $\mathbb{C}_n^\omega(A)$ we infer that the complements of languages in $\mathbb{C}_n(A)$ are also non-ambiguous context free $\omega$-languages. We conclude the study of classes $\mathbb{C}_n(A)$ by showing that they are neither closed under union nor under intersection.

For all previously studied decidable winning conditions for pushdown games the set of winning positions for any player had been shown to be regular. In [19, 20] Serre proved that every deterministic context free language may occur as a winning set for Eve in a pushdown game equipped with a winning condition in the form $\Omega_B$, where $B$ is a deterministic pushdown automaton. The exact nature of these sets remains open and the question is raised in [19, 20] whether there exists a pushdown game equipped with a winning condition in the form $\Omega_{A_1 \triangleright \ldots \triangleright A_n \triangleright A_{n+1}}$ such that the set of winning positions for Eve is not a deterministic context free language. We give a positive answer to this question, giving examples of winning sets which are non-deterministic non-ambiguous context free languages, or inherently ambiguous context free languages, or even non context free languages.

The paper is organized as follows. In section 2 we recall definitions and results about pushdown automata, context free ($\omega$)-languages, pushdown games, and winning conditions in the form $\Omega_{A_1 \triangleright \ldots \triangleright A_n \triangleright A_{n+1}}$. In section 3 are studied the classes $\mathbb{C}_n(A)$. Results on sets of winning positions are presented in Section 4.
2. Recall of previous definitions and results

2.1. Pushdown automata

We assume the reader to be familiar with the theory of formal (ω)-languages [22, 21, 12]. We shall use usual notations of formal language theory.

When A is a finite alphabet, a non-empty finite word over A is any sequence $x = a_1 \ldots a_k$, where $a_i \in A$ for $i = 1, \ldots, k$, and $k$ is an integer $\geq 1$. The length of $x$ is $k$, denoted by $|x|$. The empty word has no letter and is denoted by $\lambda$; its length is 0. For $x = a_1 \ldots a_k$, we write $x(i) = a_i$ and $x[i] = x(1) \ldots x(i)$ for $i \leq k$ and $x[0] = \lambda$. $A^*$ is the set of finite words (including the empty word) over $A$ and $A^+ = A^* - \{\lambda\}$.

The first infinite ordinal is $\omega$. An $\omega$-word over $A$ is an $\omega$-sequence $a_1 \ldots a_n \ldots$, where for all integers $i \geq 1$, $a_i \in A$. When $\sigma$ is an $\omega$-word over $A$, we write $\sigma = \sigma(1)\sigma(2)\ldots\sigma(n)\ldots$, where for all $i$, $\sigma(i) \in A$, and $\sigma[n] = \sigma(1)\sigma(2)\ldots\sigma(n)$ for all $n \geq 1$ and $\sigma[0] = \lambda$.

The prefix relation is denoted $\subseteq$: a finite word $u$ is a prefix of a finite word $v$ (respectively, an infinite word $v$), denoted $u \subseteq v$, if and only if there exists a finite word $w$ (respectively, an infinite word $w$), such that $v = u.w$. The set of $\omega$-words over the alphabet $A$ is denoted by $A^{\omega}$. An $\omega$-language over an alphabet $A$ is a subset of $A^{\omega}$.

In [19, 20] deterministic pushdown automata are defined with two restrictions. It is supposed that there are no $\lambda$-transitions, i.e. the automata are real time. Moreover one can push at most one symbol in the pushdown stack using a single transition of the automaton.

We now define pushdown automata, keeping this second restriction but allowing the existence of $\lambda$-transitions; and we define also the non deterministic version of pushdown automata.

A pushdown automaton (PDA) is a 6-tuple $A = (Q, \Gamma, A, \bot, q_{in}, \delta)$, where $Q$ is a finite set of states, $\Gamma$ is a finite pushdown alphabet, $A$ is a finite input alphabet, $\bot$ is the bottom of stack symbol, $q_{in} \in Q$ is the initial state, and $\delta$ is the transition relation which is a mapping from $Q \times (A \cup \{\lambda\}) \times \Gamma$ to subsets of

$$\{\text{skip}(q), \text{pop}(q), \text{push}(q, \gamma) \mid q \in Q, \gamma \in \Gamma - \{\bot\}\}$$

The bottom symbol appears only at the bottom of the stack and is never popped thus for all $q, q' \in Q$ and $a \in A$, it holds that $\text{pop}(q') \notin \delta(q, a, \bot)$.

The pushdown automaton $A$ is deterministic if for all $q \in Q$, $a \in A$ and $Z \in \Gamma$, the set $\delta(q, a, Z)$ contains at most one element; moreover if for some $q \in Q$ and $Z \in \Gamma$, $\delta(q, \lambda, Z)$ is non-empty then for all $a \in A$ the set $\delta(q, a, Z)$ is empty.

If $\sigma \in \Gamma^+$ describes the pushdown store content, the rightmost symbol will be assumed to be on “top” of the store. A configuration of the pushdown automaton $A$ is a pair $(q, \sigma)$ where $q \in Q$ and $\sigma \in \Gamma^*$.

For $a \in A \cup \{\lambda\}$, $\sigma \in \Gamma^*$ and $Z \in \Gamma$:

if $(\text{skip}(q'))$ is in $\delta(q, a, Z)$, then we write $a : (q, \sigma, Z) \rightarrow_{A} (q', \sigma, Z)$;
if \((\text{pop}(q'))\) is in \(\delta(q, a, Z)\), then we write \(a : (q, \sigma, Z) \rightarrow_A (q', \sigma)\);
if \((\text{push}(q', \gamma))\) is in \(\delta(q, a, Z)\), then we write \(a : (q, \sigma, Z) \rightarrow_A (q', \sigma, Z, \gamma)\).

\(\rightarrow_A^\ast\) is the transitive and reflexive closure of \(\rightarrow_A\). (The subscript \(A\) will be omitted whenever the meaning remains clear).

Let \(x = a_1a_2\ldots a_n\) be a finite word over \(A\). A finite sequence of configurations \(r = (q_i, \gamma_i)_{1\leq i \leq p}\) is called a run of \(A\) on \(x\), starting in configuration \((q, \gamma)\), iff:

1. \((q_1, \gamma_1) = (q, \gamma)\)
2. for each \(i, 1 \leq i \leq (p - 1)\), there exists \(b_i \in A \cup \{\lambda\}\) satisfying \(b_i : (q_i, \gamma_i) \rightarrow_A (q_{i+1}, \gamma_{i+1})\)
3. \(a_1a_2\ldots a_n = b_1b_2\ldots b_{p-1}\)

A run \(r\) of \(A\) on \(x\), starting in configuration \((q_{in}, \perp)\), will be simply called “a run of \(A\) on \(x\)”.

Let \(x = a_1a_2\ldots a_n\ldots\) be an \(\omega\)-word over \(A\). An infinite sequence of configurations \(r = (q_i, \gamma_i)_{i\geq 1}\) is called a run of \(A\) on \(x\), starting in configuration \((q, \gamma)\), iff:

1. \((q_1, \gamma_1) = (q, \gamma)\)
2. for each \(i \geq 1\), there exists \(b_i \in A \cup \{\lambda\}\) satisfying \(b_i : (q_i, \gamma_i) \rightarrow_A (q_{i+1}, \gamma_{i+1})\)
3. either \(a_1a_2\ldots a_n\ldots = b_1b_2\ldots b_n\ldots\)
   or \(b_1b_2\ldots b_n\ldots\) is a finite prefix of \(a_1a_2\ldots a_n\ldots\)

The run \(r\) is said to be complete when \(a_1a_2\ldots a_n\ldots = b_1b_2\ldots b_n\ldots\)
A complete run \(r\) of \(A\) on \(x\), starting in configuration \((q_{in}, \perp)\), will be simply called “a run of \(A\) on \(x\)”.

If the pushdown automaton \(A\) is equipped with a set of final states \(F \subseteq Q\),
the finitary language \(\text{accepted by } (A, F)\) is:

\[
L_f^f (A, F) = \{ x \in A^\ast \mid \text{there exists a run } r = (q_i, \gamma_i)_{1\leq i \leq p} \text{ of } A \text{ on } x \text{ such that } q_p \in F \}
\]

The class \(\text{CFL} \) of context free languages is the class of finitary languages which are accepted by pushdown automata by final states.
Notice that other accepting conditions by PDA have been shown to be equivalent to the acceptance condition by final states. Let us cite, [1]: (a) acceptance by empty storage, (b) acceptance by final states and empty storage, (c) acceptance by topmost stack letter, (d) acceptance by final states and topmost stack letter.
The class \(\text{DCFL} \) of deterministic context free languages is the class of finitary languages which are accepted by deterministic pushdown automata (DPDA) by final states.
Notice that for DPDA, acceptance by final states is not equivalent to acceptance by empty storage: this is due to the fact that a language accepted by a DPDA by empty storage must be \textit{prefix-free} while this is not necessary in the case of acceptance by final states [1].
The \( \omega \)-language \textit{B"uchi accepted by} \((A, F)\) is:

\[
L(A, F) = \{ x \in A^\omega \mid \text{there exists a run } r \text{ of } A \text{ on } x \text{ such that } In(r) \cap F \neq \emptyset \}
\]

where \( In(r) \) is the set of all states entered infinitely often during run \( r \).

If instead the pushdown automaton \( A \) is equipped with a set of accepting sets of states \( F \subseteq 2^Q \), the \( \omega \)-language \textit{Muller accepted by} \((A, F)\) is:

\[
L(A, F) = \{ x \in A^\omega \mid \text{there exists a run } r \text{ of } A \text{ on } x \text{ such that } In(r) \in F \}
\]

The class \( CF\ell_\omega \) of \textit{context free} \( \omega \)-languages is the class of \( \omega \)-languages which are B"uchi or Muller accepted by pushdown automata.

Another usual acceptance condition for \( \omega \)-words is the parity condition. In that case a pushdown automaton \( A = (Q, \Gamma, A, \bot, q_{in}, \delta) \) is equipped with a function \( col \) from \( Q \) to a finite set of colors \( C \subset \mathbb{N} \). The \( \omega \)-language accepted by \((A, col)\) is:

\[
L(A, col) = \{ x \in A^\omega \mid \text{there exists a run } r \text{ of } A \text{ on } x \text{ such that } sc(r) \text{ is even} \}
\]

where \( sc(r) \) is the smallest color appearing infinitely often in the run \( r \).

It is easy to see that a B"uchi acceptance condition can be expressed as a parity acceptance condition which itself can be expressed as a Muller condition. Thus the class of \( \omega \)-languages which are accepted by pushdown automata with a parity acceptance condition is still the class \( CF\ell_\omega \).

Consider now \textit{deterministic} pushdown automata. If \( A \) is a deterministic pushdown automaton, then for every \( \sigma \in A^\omega \), there exists at most one run \( r \) of \( A \) on \( \sigma \) determined by the starting configuration. The pushdown automaton has the continuity property iff for every \( \sigma \in A^\omega \), there exists a unique run of \( A \) on \( \sigma \) and this run is complete. It is shown in [8] that each \( \omega \)-language accepted by a deterministic B"uchi (respectively, Muller) pushdown automaton can be accepted by a deterministic B"uchi (respectively, Muller) pushdown automaton with the continuity property. The same proof works in the case of deterministic pushdown automata with parity acceptance condition.

The class of \( \omega \)-languages accepted by deterministic B"uchi pushdown automata is a strict subclass of the class \( DCF\ell\omega \) of \( \omega \)-languages accepted by deterministic pushdown automata with a Muller condition.

One can easily show that \( DCF\ell\omega \) is also the class of \( \omega \)-languages accepted by DPDA with a parity acceptance condition.

Each \( \omega \)-language in \( DCF\ell\omega \) can be accepted by a deterministic pushdown automaton \textit{having the continuity property} with parity (or Muller) acceptance condition. One can then show that the class \( DCF\ell\omega \) is closed under complementation.

The notion of ambiguity for context free \( \omega \)-languages has been firstly studied in [10]. A context free \( \omega \)-language is non ambiguous iff it is accepted by a B"uchi or Muller pushdown automaton such that every \( \omega \)-word on the input alphabet has at most one accepting run. Notice that we consider here that two runs
are equal iff they go through the same infinite sequence of configurations and \( \lambda \)-transitions occur at the same steps of the computations.

The class \( NA - CFL_\omega \) is the class of non ambiguous context free \( \omega \)-languages.

The inclusion \( DCFL_\omega \subseteq NA - CFL_\omega \) will be useful in the sequel. We shall denote \( Co - NA - CFL_\omega \) the class of complements of non ambiguous context free \( \omega \)-languages.

### 2.2. Pushdown games

Recall first that a pushdown process may be viewed as a PDA without input alphabet and initial state. A pushdown process is a 4-tuple \( P = (Q, \Gamma, \bot, \delta) \), where \( Q \) is a finite set of states, \( \Gamma \) is a finite pushdown alphabet, \( \bot \) is the bottom of stack symbol, and \( \delta \) is the transition relation which is a mapping from \( Q \times \Gamma \) to subsets of

\[
\{ \text{skip}(q), \text{pop}(q), \text{push}(q, \gamma) \mid q \in Q, \gamma \in \Gamma - \{\bot\} \}
\]

Configurations of a pushdown process are defined as for PDA. A configuration of the pushdown process \( P \) is a pair \( (q, \sigma) \) where \( q \in Q \) and \( \sigma \in \Gamma^* \).

To a pushdown process \( P = (Q, \Gamma, \bot, \delta) \) is naturally associated a pushdown graph \( G = (V, \rightarrow) \) which is a directed graph. The set of vertices \( V \) is the set of configurations of \( P \). The edge relation \( \rightarrow \) is defined as follows: \( (q, \sigma) \rightarrow (q', \sigma') \) iff the configuration \( (q', \sigma') \) can be reached in one transition of \( P \) from the configuration \( (q, \sigma) \).

We shall consider in the sequel infinite games between two players named Eve and Adam on such pushdown graphs.

So we shall assume that the set \( Q \) of states of a pushdown process is partitioned in two sets \( Q_E \) and \( Q_A \). A configuration \( (q, \sigma) \) is in \( V_E \) iff \( q \) is in \( Q_E \) and it is in \( V_A \) iff \( q \) is in \( Q_A \) so \((V_E, V_A)\) is a partition of the set of configurations \( V \).

The game graph \((V_E, V_A, \rightarrow)\) is called a pushdown game graph.

A play from a vertex \( v_1 \) of this graph is defined as follows. If \( v_1 \in V_E \), Eve chooses a vertex \( v_2 \) such that \( v_1 \rightarrow v_2 \); otherwise Adam chooses such a vertex. If there is no such vertex \( v_2 \) the play stops. Otherwise the play may continue. If \( v_2 \in V_E \), Eve chooses a vertex \( v_3 \) such that \( v_2 \rightarrow v_3 \); otherwise Adam chooses such a vertex. If there is no such vertex \( v_3 \) the play stops. Otherwise the play continues in the same way.

So a play starting from the vertex \( v_1 \) is a finite or infinite sequence of vertices \( v_1 v_2 v_3 \ldots \) such that for all \( i \), \( v_i \rightarrow v_{i+1} \). We may assume, as in [19] [20], that in fact all plays are infinite.

A winning condition for Eve is a set \( \Omega \subseteq V^\omega \). An infinite two-player pushdown game is a 4-tuple \((V_E, V_A, \rightarrow, \Omega)\), where \((V_E, V_A, \rightarrow)\) is a pushdown game graph and \( \Omega \subseteq V^\omega \) is a winning condition for Eve.

In a pushdown game equipped with the winning condition \( \Omega \), Eve wins a play \( v_1 v_2 v_3 \ldots \) iff \( v_1 v_2 v_3 \ldots \in \Omega \).

A strategy for Eve is a partial function \( f : V^* \times V_E \rightarrow V \) such that, for all \( x \in V^* \) and \( v \in V_E \), \( v \rightarrow f(x, v) \).

Eve uses the strategy \( f \) in a play \( v_1 v_2 v_3 \ldots \) iff for all \( i \), \( v_i \in V_E, v_{i+1} = f(v_1 v_2 \ldots v_i) \).
A strategy \( f \) is a winning strategy for Eve from some position \( v_1 \) iff Eve wins all plays starting from \( v_1 \) and during which she uses the strategy \( f \).

A vertex \( v \in V \) is a winning position for Eve iff she has a winning strategy from it.

The notions of winning strategy and winning position are defined for the other player Adam in a similar way.

The set of winning positions for Eve and Adam will be respectively denoted by \( W_E \) and \( W_A \).

### 2.3. Winning condition \( \Omega_{A_1 \triangleright \ldots \triangleright A_n \triangleright A_{n+1}} \)

We first recall the definition of \( \omega \)-languages in the form \( L(A_1 \triangleright \ldots \triangleright A_n \triangleright A_{n+1}) \) which are used in \([19, 20]\) to define the winning conditions \( \Omega_{A_1 \triangleright \ldots \triangleright A_n \triangleright A_{n+1}} \).

We shall need the notion of limit of an infinite sequence of finite words over some finite alphabet \( A \).

Let then \((\beta_n)_{n \geq 0}\) be an infinite sequence of words \( \beta_n \in A^* \). The finite or infinite word \( \lim_{n \in \omega} \beta_n \) is determined by the set of its (finite) prefixes: for all \( v \in A^* \),

\[ v \subseteq \lim_{n \in \omega} \beta_n \iff \exists n \forall p \geq n \beta_p[|v|] = v. \]

Let now \( A = (Q, \Gamma, A, \bot, q_{in}, \delta) \) be a pushdown automaton reading words over the alphabet \( A \) and let \( \alpha \in A^* \). The pushdown stack of \( A \) is said to be strictly unbounded during a run \( r = (q_i, \gamma_i)_{i \geq 1} \) of \( A \) on \( \alpha \) iff \( \lim_{n \geq 1} \gamma_n \) is infinite.

We define now \( \omega \)-languages \( L(A_1 \triangleright \ldots \triangleright A_n \triangleright A_{n+1}) \) in a slightly more general case than in \([20]\), because this will be useful in the next section. Notice that in \([20]\), these \( \omega \)-languages are only defined in the case where \( A_1, \ldots, A_n \), are real-time deterministic pushdown automata, and \( A_{n+1} \) is a real-time deterministic pushdown automaton equipped with a parity or a Büchi acceptance condition.

Let \( n \) be an integer \( \geq 0 \) and \( A_1, A_2, \ldots A_n \), be some deterministic pushdown automata (in the case \( n = 0 \) there are not any such automata).

Let \((A_{n+1}, C)\) be a pushdown automaton equipped with a Büchi or a parity acceptance condition.

The input alphabet of \( A_1 \) is denoted \( A \) and we assume that, for each integer \( i \in [1, n] \), the input alphabet of \( A_{i+1} \) is the stack alphabet of \( A_i \).

We define inductively the \( \omega \)-language \( L(A_1 \triangleright \ldots \triangleright A_n \triangleright A_{n+1}) \subseteq A^\omega \) by:

1. If \( n = 0 \), \( L(A_1 \triangleright \ldots \triangleright A_n \triangleright A_{n+1}) = L(A_{n+1}, C) \) is the \( \omega \)-language accepted by \( A_{n+1} \) with acceptance condition \( C \).

2. If \( n > 0 \), \( L(A_1 \triangleright \ldots \triangleright A_n \triangleright A_{n+1}) \) is the set of \( \omega \)-words \( \alpha \in A^\omega \) such that:
   - When \( A_1 \) reads \( \alpha \), the stack of \( A_1 \) is strictly unbounded hence the sequence of stack contents has an infinite limit \( \alpha_1 \).
   - \( \alpha_1 \in L(A_2 \triangleright \ldots \triangleright A_n \triangleright A_{n+1}) \).

Let now \((V_E, V_A, \rightarrow)\) be a pushdown game graph associated with a pushdown process \( P \). An infinite play \( v_1v_2v_3 \ldots \), where \( v_i = (q_i, \gamma_i) \), is in the set \( \Omega_{A_1 \triangleright \ldots \triangleright A_n \triangleright A_{n+1}} \) iff:
1. The pushdown stack of $P$ is strictly unbounded during the play, i.e. $\lim_{n \geq 1} \gamma_n$ is infinite, and

2. $\lim_{n \geq 1} \gamma_n \in L(A_1 \rhd \ldots \rhd A_n \rhd A_{n+1})$.

3. Classes $C_n(A)$

3.1. Classes $C_n(A)$ and context free $\omega$-languages

For each integer $n \geq 0$ and each finite alphabet $A$ the class $C_n(A)$ is defined in [20] as the class of $\omega$-languages in the form $L(A_1 \rhd \ldots \rhd A_n \rhd A_{n+1})$, where $A_1, \ldots, A_n$ are real-time deterministic pushdown automata, the input alphabet of $A_1$ being $A$, and $A_{n+1}$ is a real-time deterministic pushdown automaton equipped with a parity acceptance condition. It is easy to see that we obtain the same class $C_n(A)$ if we restrict the definition to the case of real-time deterministic pushdown automata $A_1, \ldots, A_n, A_{n+1}$, having the continuity property. We shall denote $C_n^\lambda(A)$ the class obtained in the same way except that the deterministic pushdown automata $A_1, \ldots, A_n, A_{n+1}$, having still the continuity property, may have $\lambda$-transitions, i.e. may be non real time.

In the sequel of this paper when we consider languages in the form $L(A_1 \rhd \ldots \rhd A_n \rhd A_{n+1})$, we shall always implicitly assume that the pushdown automata $A_1, \ldots, A_n, A_{n+1}$, have the continuity property, and that, for each integer $i \in [1, n]$, the input alphabet of $A_{i+1}$ is the stack alphabet of $A_i$.

In order to prove that classes $C_n(A)$, $C_n^\lambda(A)$, are included in the class of context free $\omega$-languages we first state the following lemma.

**Lemma 3.1.** Let $A_1 = (Q_1, \Gamma_1, A_1, \bot_1, q_0^1, \delta_1)$ be a deterministic pushdown automaton and $A_2 = (Q_2, \Gamma_2, \Gamma_1, \bot_2, q_0^2, \delta_2)$ be a pushdown automaton equipped with a set of final states $F \subseteq Q_2$. Then the $\omega$-language $L(A_1 \rhd A_2)$ is a context free $\omega$-language.

**Proof.** Let $A_1 = (Q_1, \Gamma_1, A_1, \bot_1, q_0^1, \delta_1)$ be a deterministic pushdown automaton and $A_2 = (Q_2, \Gamma_2, \Gamma_1, \bot_2, q_0^2, \delta_2)$ be a pushdown automaton equipped with a set of final states $F \subseteq Q_2$.

Recall that an $\omega$-word $\alpha \in A_1^\omega$ is in $L(A_1 \rhd A_2)$ iff:

- When $A_1$ reads $\alpha$, the stack of $A_1$ is strictly unbounded hence the sequence of stack contents has an infinite limit $\alpha_1$.
- $\alpha_1 \in L(A_2, F)$.

We can decompose the reading of an $\omega$-word $\alpha \in L(A_1 \rhd A_2)$ by the pushdown automaton $A_1$ in the following way.

When reading $\alpha$, $A_1$ goes through the infinite sequence of configurations $(q_i, \gamma_i)_{i \geq 1}$. The infinite sequence of stack contents $(\gamma_i)_{i \geq 1}$ has limit $\alpha_1$ thus for each integer $j \geq 1$, there is a smallest integer $n_j$ such that, for all integers $i \geq n_j$, $\alpha_1[j] = \gamma_i[j]$. 
Recall that Cohen and Gold proved in [7] that if context free.

If the substitution is not and we consider the substitution

\[ \star \]

\[ f \] (\[ L \]) + 1\[ \rightarrow \] \[ A \] [2].

We define now a new alphabet

\[ \alpha \] = \[ \sigma_1 \sigma_2 \ldots \sigma_n \ldots \]

where for all integers \( j \geq 1 \), \( \sigma_j \in A_1^* \) and

\[ \sigma_j: (q_{n_j}, \alpha_1[j]) \rightarrow_{A_1} (q_{n_{j+1}}, \alpha_1[j+1]) = (q_{n_{j+1}}, \alpha_1[j]. \alpha_1(j+1)) \]

Notice that \( n_1 = 1 \), \( q_1 = q_0 \) and \( \alpha_1[1] = \bot_1 \) hence \( \sigma_1: (q_0^1, \bot_1) \rightarrow_{A_1} (q_{n_2}, \alpha_1[2]) \).

Let now, for each \( q, q' \in Q_1 \) and \( a, b \in \Gamma_1 \), the language \( L(q, q', a, b) \) be the set of words \( \sigma \in A_1^* \) such that:

\[ \sigma: (q, a) \rightarrow_{A_1} (q', a, b) \].

This language of finite words over \( A_1 \) is accepted by the pushdown automaton \( A_1 \) with the following modifications: the initial configuration is \( (q, a) \) and the acceptance is by final state \( q' \) and by final stack content \( a, b \). It is easy to see that this language is also accepted by a deterministic pushdown automaton by final states so it is in the class \( DCFL \).

Then each word \( \sigma_j \) belongs to the deterministic context free language

\[ L(q_{n_j}, q_{n_{j+1}}, \alpha_1[j], \alpha_1(j+1)) = \{ \sigma \in A_1^* \mid \sigma: ((q_{n_j}, \alpha_1[j]) \rightarrow_{A_1} (q_{n_{j+1}}, \alpha_1[j+1]) \}\]

In order to describe the \( \omega \)-language \( L(A_1 \triangleright A_2) \) from the \( \omega \)-language \( L(A_2, F) \) and the deterministic context free languages \( L(q, q', a, b) \) for \( q, q' \in Q_1 \) and \( a, b \in \Gamma_1 \), we now recall the notion of substitution.

A \textit{substitution} is a mapping \( f : \Sigma \rightarrow 2^{\Gamma^*} \), where \( \Sigma \) and \( \Gamma \) are two finite alphabets. If \( \Sigma = \{ a_1, \ldots, a_n \} \), then for all integers \( i \in [1; n] \), \( f(a_i) = L_i \) is a finitary language over the alphabet \( \Gamma \).

Now this mapping is extended in the usual manner to finite words: for all letters \( a_{i_1}, \ldots, a_{i_n} \in \Sigma \),

\( f(a_{i_1} \ldots a_{i_n}) = f(a_{i_1}) \ldots f(a_{i_n}) \), and to finitary languages \( \subseteq \Sigma^* \): \( f(L) = \cup_{x \in L} f(x) \).

If for each letter \( a \in \Sigma \), the language \( f(a) \) does not contain the empty word, then the substitution is said to be \( \lambda \)-free and the mapping \( f \) may be extended to \( \omega \)-words:

\( f(x(1) \ldots x(n) \ldots) = \{ u_1 \ldots u_n \ldots \mid \forall i \geq 1 \quad u_i \in f(x(i)) \} \)

and to \( \omega \)-languages \( \subseteq \Sigma^\omega \) by setting \( f(L) = \cup_{x \in L} f(x) \subseteq \Gamma^\omega \).

If the substitution is not \( \lambda \)-free we can define \( f(L) \) in the same way for \( \subseteq \Sigma^\omega \) but this time \( f(L) \subseteq \Gamma^* \cup \Gamma^\omega \), i.e. \( f(L) \) may contain finite or \( \omega \) words.

The substitution \( f \) is said to be a context free substitution iff for all \( a \in \Sigma \) the finitary language \( f(a) \) is context free.

Recall that Cohen and Gold proved in [7] that if \( L \) is a context free \( \omega \)-language and \( f \) is a context free substitution then \( f(L) \cap \Gamma^\omega \) and \( f(L) \cap \Gamma^\omega \) are context free.

We define now a new alphabet

\( \Delta = \{ L(q, q', a, b) \mid q, q' \in Q_1 \text{ and } a, b \in \Gamma_1 \} \)

and we consider the substitution \( h : \Gamma_1 \rightarrow 2^\Delta \) defined, for all \( b \in \Gamma_1 \), by:

\( h(b) = \{ L(q, q', a, b) \mid q, q' \in Q_1 \text{ and } a \in \Gamma_1 \} \)
Applying this substitution to the \( \omega \)-language \( L(A_2, F) \subseteq \Gamma_1^\omega \), we get \( h(L(A_2, F)) \). The substitution \( h \) is \( \lambda \)-free thus \( h(L(A_2, F)) \) is a \( \omega \)-language over \( \Delta \). Moreover for each \( b \in \Gamma_1 \) the set \( h(b) \) is finite hence context free. Thus \( h(L(A_2, F)) \subseteq \Delta^\omega \) is a context free \( \omega \)-language because \( L(A_2, F) \) is a context free \( \omega \)-language and the substitution \( h \) is a context free substitution.

Let now \( R \subseteq \Delta^\omega \) be the \( \omega \)-language defined as follows. An \( \omega \)-word \( x \in R \) has its first letter in the set \( \{ L(q_0^1, q', \bot_1, b) \mid q' \in Q_1 \text{ and } b \in \Gamma_1 \} \), and each letter \( L(q, q', a, b) \), for \( q, q' \in Q_1 \) and \( a, b \in \Gamma_1 \), in \( x \) is followed by a letter in the set \( \{ L(q', q'', b, c) \mid q'' \in Q_1 \text{ and } c \in \Gamma_1 \} \).

The \( \omega \)-language \( R \) is regular thus \( h(L(A_2, F)) \cap R \subseteq \Delta^\omega \) is a context free \( \omega \)-language because the class \( CFL_\omega \) is closed under intersection with regular \( \omega \)-languages \(^7\).

Consider now the substitution \( \Theta : \Delta \to 2^{\Delta_1^\omega} \) defined, for all letters \( L(q, q', a, b) \in \Delta \), by \( \Theta(L(q, q', a, b)) = L_{(q, q', a, b)} \). The substitution \( \Theta \) is context free thus

\[
\Theta[h(L(A_2, F)) \cap R] \cap \Delta_1^\omega
\]

is a context free \( \omega \)-language and so is \( \bot_1.( \Theta[h(L(A_2, F)) \cap R] \cap \Delta_1^\omega ) \). By construction this \( \omega \)-language is \( L(A_1 \triangleright A_2) \).

We can in fact obtain a refined result if the language \( L(A_2, F) \) is non ambiguous.

**Lemma 3.2.** Let \( A_1 = (Q_1, \Gamma_1, A_1, \bot_1, q_0^1, \delta_1) \) be a deterministic pushdown automaton and \( A_2 = (Q_2, \Gamma_2, \Gamma_1, \bot_2, q_0^2, \delta_2) \) be a pushdown automaton equipped with a set of final states \( F \subseteq Q_2 \). If the \( \omega \)-language \( L(A_2, F) \) is non ambiguous then \( L(A_1 \triangleright A_2) \in NA - CFL_\omega \).

**Proof.** Let \( A_1 = (Q_1, \Gamma_1, A_1, \bot_1, q_0^1, \delta_1) \) be a deterministic pushdown automaton and \( A_2 = (Q_2, \Gamma_2, \Gamma_1, \bot_2, q_0^2, \delta_2) \) be a pushdown automaton equipped with a set of final states \( F \subseteq Q_2 \).

We assume that \( L(A_2, F) \) is non ambiguous so we can assume, without loss of generality, that the pushdown automaton \( A_2 \) itself is non ambiguous.

We are going to explain informally the construction of a non ambiguous Büchi pushdown automaton \( A \) accepting the \( \omega \)-language \( L(A_1 \triangleright A_2) \).

We refer now to the proof of the preceding lemma. We have considered the reading of an \( \omega \)-word \( \alpha \in L(A_1 \triangleright A_2) \) by \( A_1 \), and we have shown that the word \( \alpha \) can then be decomposed in the form

\[
\alpha = \sigma_1.\sigma_2 \ldots \sigma_n \ldots
\]

where for all integers \( j \geq 1 \), \( \sigma_j \) belongs to the deterministic context free language

\[
L_{(q_n, q_{n+1}, \alpha_1(j), \alpha_1(j+1))} = \{ \sigma \in A_1^* \mid (q_{n+1}, \sigma) \mapsto_{A_1} (q_n, \alpha_1(j)+1) \}
\]

We can see that the integers \( n_j \) were defined in a unique way. However there may exist several decompositions of the \( \omega \)-word \( \alpha \) into words of languages \( L_{(q, q', a, b)} \).

In order to ensure a unique decomposition we are going to slightly modify the definition of these languages.

For each \( q, q' \in Q_1 \) and \( a, b \in \Gamma_1 \), the language \( U_{(q, q', a, b)} \) is the set of words \( \sigma \in A_1^* \) such that:
(a) $\sigma : (q, a) \mapsto_{A_1}^* (q', a.b)$ and

(b) If for some $\sigma' \supseteq \sigma$ and $s \in Q$, $\sigma' : (q, a) \mapsto_{A_1}^* (s, a.b)$ then there is a word $u \in A_1^*$ and a state $t \in Q$, such that $\sigma'.u \subseteq \sigma$ and $u : (s, a.b) \mapsto_{A_1}^* (t, a)$.

(c) If there is a run $(q_i, \gamma_i)_{1 \leq i \leq p}$ of $A_1$ on $\sigma$ such that $(q_1, \gamma_1) = (q, a)$ and $(q_p, \gamma_p) = (s, a.b)$ for some $s \in Q$, $s \neq q'$, then either there is an integer $p' < p$ such that $(q_i, \gamma_i)_{1 \leq i \leq p'}$ is a run of $A_1$ on $\sigma$ and $(q_{p'}, \gamma_{p'}) = (q', a.b)$ or it holds that $\lambda : (s, a.b) \mapsto_{A_1}^* (s', a)$ for some $s' \in Q$ and $\lambda : (s', a) \mapsto_{A_1}^* (q', a.b)$.

It is easy to see that the languages $U_{(q,q',a,b)}$ are also in the class $DCFL$ and that, for each $q, q' \in Q_1$ and $a, b \in \Gamma_1$, it holds that $U_{(q,q',a,b)} \subseteq L_{(q,q',a,b)}$.

We can see that, in the above decomposition $\alpha = \sigma_1, \sigma_2 \ldots \sigma_n \ldots$ of the $\omega$-word $\alpha$, for all integers $j \geq 1$, the word $\sigma_j$ belongs in fact to the deterministic context free language $U_{(q_{n_j}, q_{n_{j+1}}, \alpha_{1(j)}, \alpha_{1}(j+1))}$.

The rest of the proof of Lemma 3.1 can be pursued, replacing languages $L_{(q,q',a,b)}$ by languages $U_{(q,q',a,b)}$.

But now we have a unique decomposition of $\alpha$ in the form

$$\alpha = \sigma_1', \sigma_2' \ldots \sigma_n' \ldots$$

where for all integers $j \geq 1$, the word $\sigma_j'$ belongs to some language $U_{(s_j, t_j, a_j, b_j)}$ satisfying: (1) $s_1 = q_0^1$, $a_1 = \bot_1$, (2) for all integers $j \geq 1$, $t_j = s_{j+1}$ and $b_j = a_{j+1}$.

This unique decomposition is crucial in the construction of the non ambiguous Büchi PDA $A$ accepting $L(A_1 \triangleright A_2)$. We shall explain informally the behaviour of this automaton.

For each $q, q' \in Q_1$ and $a, b \in \Gamma_1$, the language $U_{(q,q',a,b)}$ is accepted by a deterministic pushdown automaton $B_{(q,q',a,b)}$ whose stack alphabet is denoted $\Gamma_{(q,q',a,b)}$. We can assume that all these alphabets are disjoint and that they are also disjoint from $\Gamma_1$, the stack alphabet of $A_1$. The stack alphabet of $A$ will be

$$\Gamma^A = \Gamma_1 \cup \bigcup_{q,q' \in Q_1 \text{ and } a,b \in \Gamma_1} \Gamma_{(q,q',a,b)}$$

When reading an $\omega$-word $\alpha \in L(A_1 \triangleright A_2)$ the pushdown automaton $A$ will guess, using the non determinism, the unique decomposition of $\alpha$ in the form

$$\alpha = \sigma_1', \sigma_2' \ldots \sigma_n' \ldots$$

where for all integers $j \geq 1$, the word $\sigma_j'$ belongs to some language $U_{(s_j, t_j, a_j, b_j)}$ satisfying: (1) $s_1 = q_0^1$, $a_1 = \bot_1$, (2) for all integers $j \geq 1$, $t_j = s_{j+1}$ and $b_j = a_{j+1}$.

In addition $A$ will simulate the reading of the $\omega$-word $\alpha_1 = a_1a_2a_3\ldots$ by the PDA $A_2$.

During a run of $A$ the stack content is always a word in the form $\bot.u.v$ where $\bot$ is the bottom symbol of $A$, $u \in (\Gamma_1 - \{\bot\})^*$ and $v$ is in $(\Gamma_{(q,q',a,b)})^*$ for some $q, q' \in Q_1$ and $a, b \in \Gamma_1$.

After having read the initial segment $\sigma_1', \sigma_2' \ldots \sigma_j'$ of $\alpha$, the content of the stack of $A$ is equal to the content of the stack of $A_2$ after having read $a_1a_2\ldots a_j$. 

\begin{itemize}
  \item \textbf{(a)} $\sigma : (q, a) \mapsto_{A_1}^* (q', a.b)$ and
  \item \textbf{(b)} If for some $\sigma' \supseteq \sigma$ and $s \in Q$, $\sigma' : (q, a) \mapsto_{A_1}^* (s, a.b)$ then there is a word $u \in A_1^*$ and a state $t \in Q$, such that $\sigma'.u \subseteq \sigma$ and $u : (s, a.b) \mapsto_{A_1}^* (t, a)$.
  \item \textbf{(c)} If there is a run $(q_i, \gamma_i)_{1 \leq i \leq p}$ of $A_1$ on $\sigma$ such that $(q_1, \gamma_1) = (q, a)$ and $(q_p, \gamma_p) = (s, a.b)$ for some $s \in Q$, $s \neq q'$, then either there is an integer $p' < p$ such that $(q_i, \gamma_i)_{1 \leq i \leq p'}$ is a run of $A_1$ on $\sigma$ and $(q_{p'}, \gamma_{p'}) = (q', a.b)$ or it holds that $\lambda : (s, a.b) \mapsto_{A_1}^* (s', a)$ for some $s' \in Q$ and $\lambda : (s', a) \mapsto_{A_1}^* (q', a.b)$.
\end{itemize}
Then $\mathcal{A}$ guesses that the next word in the decomposition of $\alpha$ belongs to some $U(n_{j+1}, t_{j+1}, a_{j+1}, b_{j+1})$. It uses the stack alphabet $\Gamma$ on the top of the stack to simulate the reading of $\sigma_{j+1}$ by $\mathcal{B}(n_{j+1}, t_{j+1}, a_{j+1}, b_{j+1})$. Then when it has guessed that it has completely read the word $\sigma_{j+1}$, it erases letters of $\Gamma(n_{j+1}, t_{j+1}, a_{j+1}, b_{j+1})$ from the stack, and simulates the reading of the letter $a_{j+1}$ by $\mathcal{A}_2$, and so on. A Büchi acceptance condition is then used to simulate the acceptance of $\alpha_1$ by $\mathcal{A}_2$.

The Büchi PDA $(A_2, F)$ is non ambiguous and the above cited decomposition of $\alpha$ is unique so there is a unique accepting run of the Büchi PDA $\mathcal{A}$ on $\alpha$.

Finally we have proved that $L(A_1 \triangleright A_2) \subseteq \text{NA} - \text{CFL}_\omega$. \hfill $\Box$

**Proposition 3.3.** Let $n$ be an integer $\geq 1$, $A_1, A_2, \ldots, A_n$, be some deterministic pushdown automata and $(A_{n+1}, C)$ be a pushdown automaton equipped with a Büchi acceptance condition. The input alphabet of $A_1$ is denoted $A$ and we assume that, for each integer $i \in [1, n]$, the input alphabet of $A_{i+1}$ is the stack alphabet of $A_i$. Then $L(A_1 \triangleright \ldots \triangleright A_n \triangleright A_{n+1}) \subseteq \text{CFL}_\omega$. Moreover if $L(A_{n+1}, C)$ is non ambiguous then $L(A_1 \triangleright \ldots \triangleright A_n \triangleright A_{n+1}) \subseteq \text{NA} - \text{CFL}_\omega$.

**Proof.** We reason by induction on the integer $n$.

For $n = 1$ the result is stated in the above Lemmas 3.1 and 3.2.

Assume now that the result is true for some integer $n \geq 1$.

Let $A_1, A_2, \ldots, A_n, A_{n+1}$, be some deterministic pushdown automata and $(A_{n+2}, C)$ be a pushdown automaton equipped with a Büchi acceptance condition such that the language $L(A_1 \triangleright \ldots \triangleright A_{n+1} \triangleright A_{n+2}) \subseteq A^\omega$ is well defined.

By induction hypothesis the language $L(A_2 \triangleright \ldots \triangleright A_{n+1} \triangleright A_{n+2})$ is a context free $\omega$-language accepted by a Büchi pushdown automaton $(A, F)$.

But by definition of the language $L(A_1 \triangleright \ldots \triangleright A_{n+1} \triangleright A_{n+2})$ it holds that

$$L(A_1 \triangleright \ldots \triangleright A_{n+1} \triangleright A_{n+2}) = L(A_1 \triangleright A)$$

thus Lemma 3.1 implies that $L(A_1 \triangleright \ldots \triangleright A_{n+1} \triangleright A_{n+2}) \subseteq \text{CFL}_\omega$.

Assume now that $L(A_{n+1}, C)$ is non ambiguous. Reasoning as above but applying Lemma 3.2 instead of Lemma 3.1 we infer that $L(A_1 \triangleright \ldots \triangleright A_{n+1} \triangleright A_{n+2})$ is in $\text{NA} - \text{CFL}_\omega$. \hfill $\Box$

In particular, Proposition 3.3 implies the following result.

**Corollary 3.4.** For each integer $n \geq 0$, the following inclusions hold:

$$C_n^{\lambda}(A) \subseteq C_n^{\alpha}(A) \subseteq \text{NA} - \text{CFL}_\omega$$

We shall later get a stronger result (see Corollary 3.3) from the study of closure properties of classes $C_n(A), C_n^{\alpha}(A)$. 

3.2. Closure properties of classes $C_n(A)$, $C_n^\lambda(A)$

We first state the following lemma.

**Lemma 3.5.** The class $C_n^\lambda(A)$ is closed under complementation.

**Proof.** Let $A_1 = (Q_1, \Gamma_1, A_1, \bot_1, q_1^0, \delta_1)$ be a deterministic pushdown automaton and $(A_2 = (Q_2, \Gamma_2, \Gamma_1, \bot_2, q_2^0, \delta_2), col_2)$ be a deterministic pushdown automaton equipped with a parity acceptance condition. Recall that an $\omega$-word $\alpha \in A_1^\omega$ is in $L(A_1 \rhd A_2)$ iff: when $A_1$ reads $\alpha$, the stack of $A_1$ is strictly unbounded and the sequence of stack contents has an infinite limit $\alpha_1 \in L(A_2, col_2)$.

Thus an $\omega$-word $\alpha \in A_1^\omega$ is in the complement of $L(A_1 \rhd A_2)$ iff one of the two following conditions holds:

1. When $A_1$ reads $\alpha$, the stack of $A_1$ is strictly unbounded and the limit $\alpha_1$ of stack contents is in the complement of $L(A_2, col_2)$.
2. When $A_1$ reads $\alpha$, the stack of $A_1$ is not strictly unbounded.

The class $DCFL_\omega$ is closed under complementation thus the complement of $L(A_2, col_2)$ is equal to $L(A_2, col_3)$, for some deterministic pushdown automaton $A_3$ equipped with a parity acceptance condition.

The language $L(A_1 \rhd A_3)$ is the set of $\omega$-words $\alpha \in A_1^\omega$ such that, when $A_1$ reads $\alpha$, the stack of $A_1$ is strictly unbounded and the limit $\alpha_1$ of stack contents is in $L(A_3, col_3)$. So we see that, in order to get the complement of $L(A_1 \rhd A_3)$ we have to add to $L(A_1 \rhd A_3)$ the set $B$ of all $\omega$-words $\alpha \in A_1^\omega$ such that, when $A_1$ reads $\alpha$, the stack of $A_1$ is not strictly unbounded.

To do this we are going first to modify the automaton $A_1$ in such a way that, when reading $\omega$-words in $B$, the stack will be strictly unbounded.

We now explain informally the behaviour of the new pushdown automaton $A_1'$. The stack alphabet of $A_1'$ is $\Gamma_1 \cup \Gamma_1'$, where $\Gamma_1' = \{\gamma' \mid \gamma \in \Gamma_1\}$ is just a copy of $\Gamma_1$, such that $\Gamma_1 \cap \Gamma_1' = \emptyset$.

The essential idea is that $A_1'$ will simulate $A_1$ but it has the additional following behaviour. Using $\lambda$-transitions it pushes in the stack letters of $\Gamma_1'$, always keeping the information about the content of the stack of $A_1$.

More precisely, if at some step while reading an $\omega$-word $\alpha \in A_1^\omega$ by $A_1$ the stack content is a finite word $\gamma = \gamma_1, \gamma_2, \ldots, \gamma_j$, where each $\gamma_i$ is a letter of $\Gamma_1$, then the corresponding stack content of $A_1'$ will be in the form $\gamma_1^{n_1} \gamma_2^{n_2} \ldots \gamma_j^{n_j}$, where $n_1, n_2, \ldots, n_j$, are positive integers.

If when $A_1$ reads $\alpha$ the stack is strictly unbounded and the limit of the stack contents is an $\omega$-word $\alpha_1$, then when $A_1'$ reads the same word $\alpha$ its stack will be also strictly unbounded and the limit of the stack contents will be an $\omega$-word $\alpha_1'$. Moreover it will hold that $(\alpha_1'/\Gamma_1') = \alpha_1$, where $(\alpha_1'/\Gamma_1')$ is the word $\alpha_1'$ from which are removed all letters in $\Gamma_1'$.

On the other hand if when $A_1$ reads $\alpha$ the stack is not strictly unbounded, the limit of the stack contents being a finite word $\alpha_1$, then when $A_1'$ reads the same word $\alpha$ its stack will be strictly unbounded and its limit will be an $\omega$-word $\alpha_1'$ such that $(\alpha_1'/\Gamma_1') = \alpha_1$. 

Notice that the stack content of $A'_1$ will always be in the form $\bot_1.(\bot_1')^* \text{ or } u.Z.(Z')^n$ for some $u \in \bot_1.(\Gamma_1 \cup \Gamma'_1)^*$, $Z \in \Gamma_1$, $Z'$ being the copy of $Z$ in $\Gamma'_1$, and $n \geq 0$ being an integer.

The behaviour of the deterministic pushdown automaton $A'_1$, reading an $\omega$-word, will be the same as the behaviour of $A_1$ but with the following modifications.

(a) Between any two transitions of $A_1$ is added a $\lambda$-transition of $A'_1$ which simply pushes in the stack, when the topmost stack letter of $A'_1$ is $Z \in \Gamma_1$ or $Z' \in \Gamma'_1$, an additional letter $Z'$.

(b) Assume now that at some step of the reading of $\alpha$ by $A'_1$ and $A_1$, and after the execution of a $\lambda$-transition as explained in above item (a), the topmost stack letter of $A'_1$ is some letter $Z' \in \Gamma'_1$. Recall that the stack content of $A'_1$ will be in the form $\bot_1.(\bot_1')^n$ (if $Z' = \bot_1'$) or $u.Z.(Z')^n$ for some $u \in \bot_1.(\Gamma_1 \cup \Gamma'_1)^*$, $Z \in \Gamma_1$, $Z'$ being the copy of $Z$ in $\Gamma'_1$, and $n \geq 1$.
Notice that the corresponding stack content of $A_1$ will be $\bot_1$ or $(u/\Gamma'_1).Z$.
Suppose now that $A_1$ reads a letter $a \in A_1$ or executes a $\lambda$-transition.
If it pushes letter $T$ in the stack then $A'_1$ would push the same letter $T$ in its stack.
If $A_1$ would skip (its topmost stack letter being $Z$), then $A'_1$ also skips.
But if $A_1$, reading the letter $a \in A_1$ or executing a $\lambda$-transition, the topmost stack letter being $Z$, would pop the letter $Z$, then $A'_1$ pops the whole segment $Z.(Z')^n$ at the top of the stack, using $\lambda$-transitions.

Notice that we do not detail here the set of states of $A'_1$. It contains the set of states $Q_1$ of $A_1$ and is sufficiently enriched, to achieve the goal of simulating the behaviour of $A_1$, adding the modifications cited above.

Assume now that when $A_1$ reads $\alpha$ its stack is strictly unbounded and the limit of the stack contents is an $\omega$-word $\alpha_1$. Then when $A'_1$ reads the same word $\alpha$ its stack is also strictly unbounded and the limit of the stack contents will be an $\omega$-word $\alpha'_1$ such that $(\alpha'_1/\Gamma'_1) = \alpha_1$.
On the other hand if when $A_1$ reads $\alpha$ the stack is not strictly unbounded, then the limit of its stack contents is a finite word $\alpha_1 = \alpha_1(1).\alpha_1(2) \ldots \alpha_1(|\alpha_1|)$.
In that case when $A'_1$ reads the same word $\alpha$ its stack will be strictly unbounded and its limit will be an $\omega$-word $\alpha'_1$ in the form

$$\alpha'_1 = \alpha_1(1).\alpha_1(1)'^{n_1}.\alpha_1(2).\alpha_1(2)'^{n_2} \ldots (\alpha_1(|\alpha_1| - 1)'^{n_{|\alpha_1|-1}}.\alpha_1(|\alpha_1|).\alpha_1(|\alpha_1|)'^{n_{|\alpha_1|-1}}$$

for some integers $n_1, n_2, \ldots, n_{|\alpha_1|-1}$. In particular it will hold that $(\alpha'_1/\Gamma'_1) = \alpha_1$.

It is now easy to modify the pushdown automaton $A_3$ in such a way that we obtain a deterministic pushdown automaton $A'_3$ equipped with parity acceptance condition $col'_3$ such that the input alphabet of $A'_3$ is $\Gamma_1 \cup \Gamma'_1$, and an $\omega$-word $\alpha'_1 \in (\Gamma_1 \cup \Gamma'_1)^\omega$ is in $L(A_3', col'_3)$ iff $(\alpha'_1/\Gamma'_1)$ is a finite word or $(\alpha'_1/\Gamma'_1)$ is infinite and is in $L(A_3, col'_3)$.
Thus the $\omega$-language $L(A'_1 \triangleright A'_3)$ is the complement of $L(A_1 \triangleright A_2)$ and this ends the proof.

**Proposition 3.6.** For each integer $n \geq 0$, the class $C_n^\omega(A)$ is closed under complementation.
**Proof.** We now reason by induction on the integer \( n \geq 0 \).

For \( n = 0 \), \( C_0^\lambda(A) = DCFL_\omega \) is known to be closed under complementation \([21]\).

For \( n = 1 \), \( C_1^\lambda(A) \) is closed under complementation by Lemma 3.5.

Assume now that we have proved that for every positive integer \( k \leq n \) the class \( C_k^\lambda(A) \) is closed under complementation.

Let \( A_1, A_2, \ldots, A_n, A_{n+1} \), be some deterministic pushdown automata and \((A_{n+2}, \text{col})\) be a deterministic pushdown automaton equipped with a parity acceptance condition such that the language \( L(A_1 \triangleright \ldots \triangleright A_{n+1} \triangleright A_{n+2}) \subseteq A_1^\omega \) is well defined.

An \( \omega \)-word \( \alpha \in A_1^\omega \) is in the complement of \( L(A_1 \triangleright \ldots \triangleright A_n \triangleright A_{n+1} \triangleright A_{n+2}) \) iff one the two following conditions holds:

1. When \( A_1 \) reads \( \alpha \), the stack of \( A_1 \) is strictly unbounded and the limit \( \alpha_1 \) of stack contents is in the complement of \( L(A_2 \triangleright \ldots \triangleright A_n \triangleright A_{n+1} \triangleright A_{n+2}) \).

2. When \( A_1 \) reads \( \alpha \), the stack of \( A_1 \) is not strictly unbounded.

By induction hypothesis the complement of the \( \omega \)-language \( L(A_2 \triangleright \ldots \triangleright A_n \triangleright A_{n+1} \triangleright A_{n+2}) \) is in \( C_n^\lambda(A) \) so it is in the form \( L(A_2' \triangleright \ldots \triangleright A_{n+1}' \triangleright A_{n+2}') \).

We can do similar modifications as in the case \( n = 1 \), replacing \( A_1 \), whose stack alphabet is \( \Gamma_1 \), by another deterministic pushdown automaton \( A_1' \), whose alphabet is \( \Gamma_1 \cup \Gamma_1' \) where \( \Gamma_1' \) is a copy of \( \Gamma_1 \).

If when \( A_1 \) reads \( \alpha \) the limit of its stack contents is a finite or infinite word \( \alpha_1 \) then when \( A_1' \) reads the same word \( \alpha \) the limit of its stack contents is an \( \omega \)-word \( \alpha_1' \) such that \( (\alpha_1'/\Gamma_1') = \alpha_1 \).

It is now easy to modify the language \( L(A_2 \triangleright \ldots \triangleright A_n \triangleright A_{n+1} \triangleright A_{n+2}) \) in such a way that we get a language \( L(A_2' \triangleright \ldots \triangleright A_n' \triangleright A_{n+2}') \) of \( \omega \)-words over \( \Gamma_1 \cup \Gamma_1' \) containing an \( \omega \)-word \( \alpha_1' \) if and only if: either \( (\alpha_1'/\Gamma_1') \) is a finite word or \( (\alpha_1'/\Gamma_1') \) belongs to the \( \omega \)-language \( L(A_2' \triangleright \ldots \triangleright A_{n+1}' \triangleright A_{n+2}') \).

Thus it holds that \( L(A_1' \triangleright \ldots \triangleright A_{n+1}' \triangleright A_{n+2}') \) is the complement of \( L(A_1 \triangleright \ldots \triangleright A_n \triangleright A_{n+1} \triangleright A_{n+2}) \).

\[ \square \]

**Remark 3.7.** In \([19, 20]\) Serre defined winning conditions \( \Omega_{A_1 \triangleright \ldots \triangleright A_n \triangleright A_{n+1}} \) for pushdown games using languages in classes \( C_n(A) \). He then showed that these winning conditions lead to decision procedures to decide the winner in pushdown games. The question now naturally arises whether the proofs can be extended to winning conditions defined in the same way from classes \( C_n^\lambda(A) \). Then the closure under complementation of these classes would be relevant from a game point of view. On the other hand this closure property provides also some more information about classes \( C_n(A) \), given by next corollary, which is already important from a game point of view.

**Corollary 3.8.** For each integer \( n \geq 0 \), the following inclusions hold:

\[ C_n(A) \subseteq C_0^\lambda(A) \subseteq NA - CFL_\omega \bigcap Col NA - CFL_\omega \]
Proof. It follows directly from Corollary 3.4 and Proposition 3.6.

We now prove that the classes $C_n(A), C^λ_n(A)$, are not closed under other boolean operations.

**Proposition 3.9.** For each integer $n \ge 0$, the classes $C_n(A)$ and $C^λ_n(A)$ are neither closed under union nor under intersection.

**Proof.** Notice first that for each integer $n \ge 0$, $C_n(A) \subseteq C_{n+1}(A)$ and $C^λ_n(A) \subseteq C^λ_{n+1}(A)$.

The $\omega$-languages $L_1 = \{a^n.b^m.c^n.d^ω \mid n, m, p \ge 1$ and $n = m\}$ and $L_2 = \{a^n.b^m.c^n.d^ω \mid n, m, p \ge 1$ and $m = p\}$, over the alphabet $A = \{a, b, c, d\}$, are in $DCFL_ω$ and they are in all classes $C_n(A)$ and $C^λ_n(A)$. But their intersection is $L_1 \cap L_2 = \{a^n.b^m.c^n.d^ω \mid n \ge 1\}$. This $\omega$-language is not context free because the finitary language $\{a^n.b^m.c^n \mid n \ge 1\}$ is not context free [11] and an $\omega$-language in the form $L.d^ω$, with $L \subseteq \{a, b, c\}^*$, is context free iff the finitary language $L$ is context free [17]. Thus $L_1 \cap L_2$ cannot be in any class $C_n(A)$ and $C^λ_n(A)$ because these classes are included in $CFL_ω$.

On the other hand consider the $\omega$-languages $L_3 = \{a^n.b^m.c^p.d^ω \mid n, m, p \ge 1$ and $n \neq m\}$ and $L_4 = \{a^n.b^m.c^p.d^ω \mid n, m, p \ge 1$ and $m \neq p\}$. These $\omega$-languages are in $DCFL_ω$ and in every class $C_n(A)$ or $C^λ_n(A)$. If the language $L_3 \cup L_4$ was in some class $C_n(A)$ or $C^λ_n(A)$, then by Proposition 3.6 its complement $L_5$ would be also in $C^λ_n(A)$ and it would be a context free $\omega$-language. This would imply that $L_5 \cap a^+.b^+.c^+.d^ω$ is context free because the class $CFL_ω$ is closed under intersection with regular $\omega$-languages. But $L_5 \cap a^+.b^+.c^+.d^ω = \{a^n.b^n.c^n.d^ω \mid n \ge 1\}$ is not context free thus for each integer $n \ge 0$, the classes $C_n(A), C^λ_n(A)$ are not closed under union.

Notice that the union $\cup_{n \ge 0} C^λ_n(A)$ is also neither closed under intersection nor under union.

### 4. Winning sets in a pushdown game

Recall that it is proved in [19] that every deterministic context free language may occur as a winning set for Eve in a pushdown game equipped with a winning condition in the form $Ω_B$, where $B$ is a deterministic pushdown automaton.

Serre asked also whether there exists a pushdown game equipped with a winning condition in the form $Ω_{A_1 > … > A_n > A_{n+1}}$ such that the set of winning positions for Eve is not a deterministic context free language.

We are going to prove in this section that such pushdown games exist, giving examples of winning sets which are non-deterministic non-ambiguous context free languages, or inherently ambiguous context free languages, or even non context free languages.

The exact form of the winning sets remains open. Serre conjectured in [18] that one could prove that, for $n \ge 0$, the winning sets for Eve in pushdown games equipped with a winning condition in the form $Ω_{A_1 > … > A_n > A_{n+1}}$, form a class of languages at level $n$, and that for $n = 0$ the winning sets could be deterministic context free languages.

So we think that, in order to better understand what is the exact form of the winning sets, it is useful to see different examples of winning sets of different complexities, and not only of the greatest complexity we have got, i.e. a non context free language.
Moreover the techniques, involving Duparc’s eraser operator, used to prove Proposition 4.3 below, are interesting by their own and are useful to understand how the games go on.

In order to present the first example we begin by recalling the operation \( x \to x^- \) which has been introduced by Duparc in his study of the Wadge hierarchy [9], where it works also on infinite words, and is also considered by Serre in [19].

For a finite word \( u \in (\Sigma \cup \{\leftarrow\})^* \), where \( \Sigma \) is a finite alphabet, the finite word \( u^- \) is inductively defined by:

\[
\lambda^+ = \lambda,
\]

and for a finite word \( u \in (\Sigma \cup \{\leftarrow\})^* \):

\[
(u.c)^- = u^- . c, \text{ if } c \in \Sigma,
\]

\[
(u. \leftarrow)^- = u^- \text{ with its last letter removed if } |u^-| > 0,
\]

i.e. \((u. \leftarrow)^- = u^- (1).u^- (2) \ldots u^- (|u^-| - 1)\) if \(|u^-| > 0\),

\[
(u. \leftarrow)^- = \lambda \text{ if } |u^-| = 0,
\]

Notice that for \( x \in (\Sigma \cup \{\leftarrow\})^* \), \( x^- \) denotes the string \( x \), once every \( \leftarrow \) occuring in \( x \), used as an eraser, has been “evaluated” to the back space operation, proceeding from left to right inside \( x \). In other words \( x^- = x \) from which every interval of the form \( “c \leftarrow” \ (c \in \Sigma) \) is removed.

For a language \( V \subseteq \Sigma^* \) we set \( V^- = \{ x \in (\Sigma \cup \{\leftarrow\})^* \mid x^- \in V \} \).

**Lemma 4.1.** Let \( L = \{ a^n . b^n \mid n \geq 1 \} \). Then \( L^- \) is a non ambiguous context free language which can not be accepted by any deterministic pushdown automaton.

**Proof.** Let \( L \) be the context free language \( \{ a^n . b^n \mid n \geq 1 \} \). The language \( L \) is a deterministic, hence non ambiguous, context free language. Thus by Theorem 6.16 of [10] the language \( L^- \) is a non ambiguous context free language.

It remains to show that \( L^- \) can not be accepted by any deterministic pushdown automaton.

The idea of the proof is essentially the same as in the proof that the context free language \( \{ a^n . b^n \mid n \geq 1 \} \cup \{ a^n . b^{2n} \mid n \geq 1 \} \) can not be accepted by any deterministic pushdown automaton. It can be found in [11] Proof of Proposition 5.3) or in [12] Exercise 6.4.4 page 251].

Towards a contradiction assume that the language \( L^- \) is accepted by a deterministic pushdown automaton \( A \). All words \( a^n . b^n \), for \( n \geq 1 \), are in the language \( L^- \). Then one could show that there exists a pair \((n, k)\), with \( n, k > 0 \), such that the accepting configurations of \( A \) reading \( a^n . b^n \) or \( a^{n+k} . b^{n+k} \) are the same. Consider now the word \( a^n . b^n, \leftarrow 2n . a . b \). It belongs to \( L^- \) and the valid computation of \( A \) reading \( a^n . b^n \) should be the beginning of the valid computation of \( A \) reading \( a^n . b^n, \leftarrow 2n . a . b \). Thus the pushdown automaton \( A \) would also accepts \( a^{n+k} . b^{n+k} . \leftarrow 2n . a . b \) which is clearly not in \( L^- \). \[\Box\]

**Lemma 4.2.** Let \( L \subseteq \Sigma^* \) be a deterministic context free language. Then there exists a pushdown process \( P = (Q, \Gamma, \bot, \delta) \), a partition \( Q = Q_E \cup Q_A \), two deterministic pushdown automata \( A_1, A_2 \), and a state \( q \in Q \) such that, in the induced pushdown game equipped with the winning condition \( \Omega_{A_1 \uplus A_2} \), one has \( \{ u \mid (q, u) \in W_E \} = L^- \).
Proof. Let \( \mathcal{P} = ((p, q), \Gamma = \Sigma \cup \{ \perp, \leftarrow , \# \}, \perp, \delta) \) be a pushdown process where \(\delta\) is defined by: \(\text{push}(p, \#) \in \delta(q, c)\) for all letters \(c \in \Sigma \cup \{ \perp, \leftarrow \}\) and \(\text{push}(p, \#) \in \delta(p, \#)\).

So the pushdown process \(\mathcal{P}\) is deterministic and its behaviour is very similar to the behaviour of the pushdown process given in the proof of Proposition 4.1 of [20]. It can only push the letter \(\#\) on the top of a given configuration.

Let \(Q = Q_E \cup Q_A\) be any partition of \(Q\).

For each configuration \((q, u, c)\), for \(c \in \Sigma \cup \{ \perp, \leftarrow \}\) and \(u \in \Gamma^*\), there is a unique infinite play starting from \((q, u, c)\), during which the pushdown stack of \(\mathcal{P}\) is strictly unbounded, and the limit of the stack contents is \(u, c, \#^\omega\).

The deterministic pushdown automaton \(A_1\) reads words over the alphabet \(\Gamma = \Sigma \cup \{ \perp, \leftarrow , \# \}\) and its stack alphabet is \(\Gamma_1 = \Sigma \cup \{ \perp \}\). Its behaviour is described as follows:

Consider first the reading of an \(\omega\)-word in the form \(\perp . u . \#^\omega\), where \(u \in (\Sigma \cup \{ \leftarrow \})^*\).

After having read the bottom symbol \(\perp\), the content of its stack is still \(\perp\). Then when the pushdown automaton \(A_1\) reads a letter \(c \in \Sigma\) it pushes the same letter in the stack. But if \(A_1\) reads the symbol \(\leftarrow\) and the topmost stack symbol is not \(\perp\) (so it is in \(\Sigma\)) then it pops the letter at the top of its stack.

So, after having read the initial segment \(\perp . u\) of \(\perp . u . \#^\omega\), the stack content of \(A_1\) is \(\perp . u^\omega\). Next the PDA \(A_1\) pushes a letter \(\#\) in the stack for each letter \(\#\) read. Thus, when \(A_1\) reads the \(\omega\)-word \(\perp . u . \#^\omega\), its stack is strictly unbounded and the limit of the stack contents is \(\perp . u^\omega . \#^\omega\).

In addition, it is easy to ensure that, when \(A_1\) reads an \(\omega\)-word which is not in \(\perp . (\Sigma \cup \{ \leftarrow \})^* . \#^\omega \cup \perp . (\Sigma \cup \{ \leftarrow \})^\omega\), then its stack is not strictly unbounded. If there is a letter \(\perp\) after the first letter of the word or if \(A_1\) reads a letter in \(\Sigma \cup \{ \leftarrow \}\) after some letter \(\#\), then the stack content remains indefinitely unchanged.

On the other hand, \(A_2\) is a deterministic pushdown automaton equipped with a parity acceptance condition which accepts the \(\omega\)-language \(\perp . L . \#^\omega\).

Consider now a given configuration \((q, \perp . u\)) of the pushdown process \(\mathcal{P}\) for some \(u \in (\Sigma \cup \{ \leftarrow , \# \})^*\), the last letter of \(u\) being not \(\#\). There is a unique infinite play starting from this position. The stack of \(\mathcal{P}\) is strictly unbounded during this play and the limit of stack contents is \(\perp . u . \#^\omega\).

When \(A_1\) reads the \(\omega\)-word \(\perp . u . \#^\omega\) its stack is strictly unbounded iff \(u \in (\Sigma \cup \{ \leftarrow \})^*\) and then the limit of stack contents is \(\perp . u^\omega . \#^\omega\).

The \(\omega\)-word \(\perp . u^\omega . \#^\omega\) is accepted by \(A_2\) iff \(u^\omega \in L\).

Thus the configuration \((q, \perp . u\)) is a winning position for Eve in the induced pushdown game, equipped with the winning condition \(\Omega_{A_1 \triangleright A_2}\), if and only if \(u \in L^\omega\). \(\square\)

We can now state the following result which follows directly from Lemmas 4.1 and 4.2.

Proposition 4.3. There exists a pushdown process \(\mathcal{P} = (Q, \Gamma, \perp, \delta)\), a partition \(Q = Q_E \cup Q_A\), two deterministic pushdown automata \(A_1, A_2\), and a state \(q \in Q\) such that, in the induced pushdown game equipped with the winning condition \(\Omega_{A_1 \triangleright A_2}\), the set \(\{u \mid (q, u) \in \mathcal{W}_E\}\) is a non-deterministic non ambiguous context free language.
Proof. Let \( L \) be the language \( \{a^n.b^n \mid n \geq 1\} \). The language \( L \) is a deterministic context free language, thus by Lemma 4.2 there exists a pushdown process \( P = (Q, \Gamma, \perp, \delta) \), a partition \( Q = Q_E \cup Q_A \), two deterministic pushdown automata \( A_1, A_2 \), and a state \( q \in Q \) such that, in the induced pushdown game equipped with the winning condition \( \Omega_{A_1 \triangleright A_2} \), one has \( \{u \mid (q, u) \in W_E\} = L^\sim \). But by Lemma 4.1 \( L^\sim \) is a non ambiguous context free language which can not be accepted by any deterministic pushdown automaton.

\[
\square
\]

Remark 4.4. In the pushdown game given in the proof of Lemma 4.2 there are some plays which are not infinite. However it is easy to find a pushdown game with the same winning set for Eve but in which all plays are infinite. The same remark will hold for pushdown games given in the proofs of the two following propositions.

We are now going to show that the set of winning positions for Eve can also be an inherently ambiguous context free language. Recall that it is well known that the language \( V = \{a^n.b^m.c^p \mid n, m, p \geq 1 \text{ and } (n = m \text{ or } m = p)\} \) is an inherently ambiguous context free language, \([11, 12]\).

Proposition 4.5. There exists a pushdown process \( P = (Q, \Gamma, \perp, \delta) \), a partition \( Q = Q_E \cup Q_A \), two deterministic pushdown automata \( A_1, A_2 \), and a state \( q \in Q \) such that, in the induced pushdown game equipped with the winning condition \( \Omega_{A_1 \triangleright A_2} \), the set \( \{u \mid (q, u) \in W_E\} \) is an inherently ambiguous context free language.

Proof. Let \( P = (\{q, q', q''\}, \Gamma = \{\perp, a, b, c, \#\}, \perp, \delta) \) be a pushdown process where \( \delta \) is defined by:

\[
\{\text{pop}(q'), \text{skip}(q'')\} \subseteq \delta(q, c), \text{pop}(q') \in \delta(q', c), \text{push}(p, \#) \in \delta(q'', c), \text{and} \\
\text{push}(p, \#) \in \delta(p, \#).
\]

We set \( Q_E = \{q\} \) and \( Q_A = \{q', q'', p\} \).

Consider now an infinite play from a given configuration \( (q, \perp.u) \), for \( u \in \{a, b, c, \#\}^* \). The topmost stack letter of this initial configuration must be a letter \( c \). Then at most two cases may happen.

1. In the first one are pushed infinitely many letters \( \# \) on the top of the stack. In this play the stack is strictly unbounded and the limit of the stack contents is \( \perp.u.\#^\omega \).

2. In the second case the letter \( c \) is popped and all next letters \( a \) are popped from the top of the stack until some letter \( b \) is on the top of the stack. From this moment infinitely many letters \( \# \) are pushed in the stack. Then the stack is strictly unbounded and the limit of the stack contents is \( \perp.a'.b.a'.b.\#^{\omega} \) if \( u = u'.b.a^k \) for some integer \( k > 0 \). Notice that this second case can only occur if \( u \) is in the form \( u = u'.b.a^k \) for some integer \( k > 0 \).

The deterministic pushdown automaton \( A_1 \) reads words over the alphabet \( \{\perp, a, b, c, \#\} \) and its stack alphabet is \( \Gamma_1 = \{\perp_1, a, b, \#\} \).

It is easy to ensure that the stack of \( A_1 \) is not strictly unbounded during the reading of an \( \omega \)-word which is not in \( W = \perp.a^\omega \cup \perp.a^+.b^\omega \cup \perp.a^+.b^+.\#^\omega \cup \perp.a^+.b+.a^+.b+.a^+.\#^\omega \).

Consider now the reading by \( A_1 \) of an \( \omega \)-word which is in \( W \). After having read the bottom symbol \( \perp \), the stack content of \( A_1 \) is still \( \perp_1 \). Then it pushes a letter \( a \) or \( b \) each time it reads the corresponding letter \( a \) or \( b \).

Then when \( A_1 \) reads an \( \omega \)-word in the form \( \perp.a^\omega \) (respectively, \( \perp.a^n.b^\omega \) for \( n \geq 1 \)) then its stack is strictly unbounded and the limit of stack contents is \( \perp_1.a^\omega \) (respectively, \( \perp_1.a^n.b^\omega \)).
If now $A_1$ reads letters $\#$ then it pushes them in the stack. In this case the input word is in the form $\bot.a^n.b^m.\#^\omega$, and the limit of stack contents of $A_1$ reading this $\omega$-word is $\bot.a^n.b^m.\#^\omega$.

If $A_1$ reads some letters $c$ after an initial segment in the form $\bot.a^n.b^m$ then it pops a letter $b$ for each letter $c$ read.

If the number of $c$ is equal to the number of $b$ of the input word, then after having read the segment $\bot.a^n.b^m.c^m$ of the input word the stack content of $A_1$ is simply $\bot.a^n$. Next $A_1$ reads the final segment $\#^\omega$ and it pushes it in the stack. So the limit of stack contents of $A_1$ reading the input $\omega$-word $\bot.a^n.b^m.c^m.\#^\omega$ is in the form $\bot.a^n.\#^\omega$

If the number of $c$ is not equal to the number of $b$ of the input word (the number of $c$ being finite or infinite), then, once this has been checked, the stack content remains unchanged so the stack will not be strictly unbounded.

One can define a deterministic pushdown automaton $A_2$, equipped with a parity acceptance condition, which accepts the $\omega$-language $\{\bot.a^n.b^n.\#^\omega \mid n \geq 1\} \cup \{ \bot.a^n.\#^\omega \mid n \geq 1\}$.

We are now going to determine the winning positions $(q, \bot.u)$ of Eve in the induced pushdown game equipped with the winning condition $\Omega_{A_1 \triangleright A_2}$.

Let $(q, \bot.u)$ be a given configuration of the pushdown process $P$ for some $u \in \{a, b, c, \#\}^*$, the last letter of $u$ being $c$. There are one or two infinite plays starting from this position. When there are two such plays, they depend on the first choice of Eve and the position $(q, \bot.u)$ is a winning position for Eve iff one of the two possible infinite plays is winning for her.

In the first play the stack is strictly unbounded and the limit of the stack contents is $\bot.u.\#^\omega$.

There is a second play if $u = u'.b.c^k$ for some integer $k > 0$. Then in this play the stack is strictly unbounded and the limit of the stack contents is $\bot.u'.b.\#^\omega$.

When $A_1$ reads the $\omega$-word $\bot.u.\#^\omega$, its stack is strictly unbounded iff $u$ is in the form $a^n.b^m.c^m$ for some $n, m \geq 1$ (the number of $c$ and of $b$ in $u$ are equal). Then the limit of stack contents is $\bot.a^n.\#^\omega$ and it is in $L(A_2)$. So $\bot.u.\#^\omega \in L(A_1 \triangleright A_2)$.

If $u = u'.b.c^k$ for some integer $k > 0$ and $A_1$ reads the $\omega$-word $\bot.u'.b.\#^\omega$ then the stack of $A_1$ is strictly unbounded iff $u'$ is in the form $a^n.b^{m-1}$ for some $n, m \geq 1$. In this case the limit of stack contents is $\bot.a^n.b^{m-1}.\#^\omega$ and it is accepted by $A_2$ iff $n = m \geq 1$.

Thus the configuration $(q, \bot.u)$ is a winning position for Eve, with the winning condition $\Omega_{A_1 \triangleright A_2}$, if and only if $u$ is in the inherently ambiguous context free language $V = \{a^n.b^m.c^p \mid n, m, p \geq 1 \text{ and } (n = m \text{ or } m = p)\}$.

**Proposition 4.6.** There exists a pushdown process $P = (Q, \Gamma, \bot, \delta)$, a partition $Q = Q_E \cup Q_A$, two deterministic pushdown automata $A_1, A_2$, and a state $q \in Q$ such that, in the induced pushdown game equipped with the winning condition $\Omega_{A_1 \triangleright A_2}$, the set $\{ u \mid (q, u) \in W_E \}$ is a non context free language.

**Proof.** We define the pushdown process $P = (Q, \Gamma, \bot, \delta)$ as in the proof of preceding Proposition 4.5 except that we set this time $Q_A = \{q\}$ and $Q_E = \{q', q''\}$, $p$. The two deterministic pushdown automata
\(A_1, A_2\), are also defined in the same way.

Consider now a configuration in the form \((q, \perp, a^n, b^m, c^p)\) for some integers \(n, m, p \geq 1\). There are two infinite plays starting from this configuration but they depend this time on the first choice of the second player Adam.

The position \((q, \perp, a^n, b^m, c^p)\) is winning for Eve iff these two infinite plays are won by her. This implies that \(n = m\) and \(m = p\).

Thus it holds that

\[
\{ u \mid (q, u) \in W_E \} \cap \perp.a^+.b^+.c^+ = \perp.\{ a^n.b^n.c^n \mid n \geq 1 \}
\]

This language is not context free because of the well known non context freeness of the language \(\{ a^n.b^n.c^n \mid n \geq 1 \}\) [1, 12].

This implies that the set \(\{ u \mid (q, u) \in W_E \}\) itself is not context free. Indeed otherwise its intersection with the rational language \(\perp.a^+.b^+.c^+\) would be context free because the class \(CFL\) is closed under intersection with rational languages.

\[ \square \]

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