An Efficient Algorithm for the Equation Tree Automaton via the k-C-Continuations

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Abstract. Champarnaud and Ziadi, and Khorsi et~al. show how to compute the equation automaton of word regular expression E via the k-C-Continuations. Kuske and Meinecke extend the computation of the equation automaton to a regular tree expression E over a ranked alphabet Σ and produce a $O(R \cdot |\mathbf{E}|^2)$ time and space complexity algorithm, where R is the maximal rank of a symbol occurring in Σ and $|\mathbf{E}|$ is the size of E. In this paper, we give a full description of the algorithm based on the acyclic minimization of Revuz. Our algorithm, which is performed in an $O(|Q| \cdot |\mathbf{E}|)$ time and space complexity, where |Q| is the number of states of the produced automaton, is more efficient than the one obtained by Kuske and Meinecke.

1 Introduction

Regular expressions, which are finite representatives of potentially infinite languages, are widely used in various application areas such as XML Schema Languages [10], logic and verification [14], etc. The concept of word regular expressions has been extended to tree regular expressions. Similarly to word expressions, one can convert them into finite recognizers, the tree automata.

The study of the different ways of conversion of regular expressions into automata and vice versa is a very active field. There exists a lot of techniques to transform regular expressions (resp. regular tree expressions) into finite automata [2,6,7,15] (resp. into finite tree automata [8,9]). As far as tree automata are concerned, computation algorithms are extensions of word cases. In [9], the computation of the position tree automaton from a regular tree expression has been achieved by extending the classical notions of Glushkov functions defined in [6], leading to the computation of an automaton which number of states is linear w.r.t. the number of occurrences of symbols but which number of transitions can be exponential. In the same paper, it is proved that this automaton can be reduced into a quadratic size recognizer.

On the other side, Kuske and Meinecke have extended the notion of word partial derivatives [1] into tree partial derivatives. They also present how to compute them extending from words to trees [8] the k-C-Continuation algorithm by Champarnaud and Ziadi [3]. They obtain an algorithm with $O(R \cdot |E| \cdot |E|)$

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space and time complexity where R is the maximal rank of a symbol occurring in the finite ranked alphabet Σ and |E| is the size of the regular expression.

In this paper, we show how to extend a notion of k-C-Continuation in order to compute from a regular tree expression its equation tree automaton with an $O(|E|+|Q|\cdot|E|)$ time and space complexity where |Q| is the number of its states. This constitutes an improvement in comparison with Kuske and Meinecke algorithm [8]. The paper is organized as follows: Section 2 outlines finite tree automata over ranked trees, regular tree expressions, and linearized regular tree expressions which allows the set of positions to be defined. Next, in Section 3 the notions of derivation and partial derivative of regular expression and set of regular expressions are introduced. Thus the definitions of equation tree automaton and k-C-Continuation tree automaton associated with the regular expression E is obtained. Afterwards, in Section 4 we present our algorithm which builds the equation tree automaton with an $O(|E|+|Q|\cdot|E|)$ time and space complexity. Finally, Section 5 provides a full example of our construction.

2 Preliminaries

Let $(\Sigma, \operatorname{ar})$ be a ranked alphabet, where Σ is a finite set and ar represents the rank of Σ which is a mapping from Σ into $\mathbb N$. The set of symbols of rank n is denoted by Σ_n . The elements of rank 0 are called constants. A tree t over Σ is inductively defined as follows: $t=a,\ t=f(t_1,\ldots,t_k)$ where a is any symbol in Σ_0 , k is any integer satisfying $k\geq 1$, f is any symbol in Σ_k and t_1,\ldots,t_k are any k trees over Σ . We denote by T_Σ the set of trees over Σ . A tree language is a subset of T_Σ . Let $\Sigma_{\geq 1}=\Sigma\backslash\Sigma_0$ denote the set of non-constant symbols of the ranked alphabet Σ . A Finite Tree Automaton (FTA) [5,8] A is a tuple (Q,Σ,Q_T,Δ) where Q is a finite set of states, $Q_T\subset Q$ is the set of final states and $\Delta\subset\bigcup_{n\geq 0}(Q\times\Sigma_n\times Q^n)$ is the set of transition rules. This set is equivalent to the function Δ from $Q^n\times\Sigma_n\to 2^Q$ defined by $(q,f,q_1,\ldots,q_n)\in\Delta\Leftrightarrow q\in\Delta(q_1,\ldots,q_n,f)$. The domain of this function can be extended to $(2^Q)^n\times\Sigma_n\to 2^Q$ as follows: $\Delta(Q_1,\ldots,Q_n,f)=\bigcup_{(q_1,\ldots,q_n)\in Q_1\times\ldots\times Q_n}\Delta(q_1,\ldots,q_n,f)$. Finally, we denote by Δ^* the function from $T_\Sigma\to 2^Q$ defined for any tree in T_Σ as follows:

 $\Delta^*(t) = \begin{cases} \Delta(a) & \text{if } t = a, a \in \Sigma_0 \\ \Delta(f, \Delta^*(t_1), \dots, \Delta^*(t_n)) & \text{if } t = f(t_1, \dots, t_n), f \in \Sigma_n, t_1, \dots, t_n \in T_\Sigma \end{cases}$ A tree is accepted by \mathcal{A} if and only if $\Delta^*(t) \cap Q_T \neq \emptyset$. The language recognized by $\mathcal{L}(\mathcal{A})$ is the set of trees accepted by \mathcal{A} i.e. $\mathcal{L}(\mathcal{A}) = \{t \in T_\Sigma \mid \Delta^*(t) \cap Q_T \neq \emptyset\}$. A state $q \in Q$ is coaccessible if $q \in Q_T$ or if $\exists Q' = \{q_1, \dots, q_n\} \subset Q, f \in \Sigma_n, q'$ a coaccessible state in Q such that $q \in Q'$ and $q' \in \Delta(f, q_1, \dots, q_n)$. The coaccessible part of the automaton \mathcal{A} is the tree automaton $\mathcal{A}' = \{Q', \Sigma, \Delta', Q_T'\}$ where $Q' = \{q \in Q \mid q \text{ is coaccessible}\}$ and $\Delta' = \{(q, f, q_1, \dots, q_n) \in \Delta \mid \{q, q_1, \dots, q_n\} \subset Q'\}$. It is easy to show that $\mathcal{L}(\mathcal{A}) = \mathcal{L}(\mathcal{A}')$.

Let \sim be an equivalence relation over Q. We denote by [q] the equivalence class of any state q in Q. The quotient of A w.r.t. \sim is the tree automaton $A_{/\sim} = (Q_{/\sim}, \Sigma, Q_{T/\sim}, \Delta_{/\sim})$ where: $Q_{/\sim} = \{[q] \mid q \in Q\}, Q_{T/\sim} = \{[q] \mid q \in Q_T\}, \Delta_{/\sim} = \{([q], f, [q_1], \dots, [q_n]) \mid (q, f, q_1, \dots, q_n) \in \Delta\}.$

For any integer $n \geq 0$, for any n languages $L_1, \ldots, L_n \subset T_{\Sigma}$, and for any symbol $f \in \Sigma_n$, $f(L_1, \ldots, L_n)$ is the tree language $\{f(t_1, \ldots, t_n) \mid t_i \in L_i\}$. The tree substitution of a constant c in Σ by a language $L \subset T_{\Sigma}$ in a tree $t \in T_{\Sigma}$, denoted by $t\{c \leftarrow L\}$, is the language inductively defined by L if t = c; $\{d\}$ if t = d where $d \in \Sigma_0 \setminus \{c\}$; $f(t_1\{c \leftarrow L\}, \ldots, t_n\{c \leftarrow L\})$ if $t = f(t_1, \ldots, t_n)$ with $f \in \Sigma_n$ and t_1, \ldots, t_n any n trees over Σ . Let c be a symbol in Σ_0 . The c-product $L_1 \cdot c L_2$ of two languages $L_1, L_2 \subset T_{\Sigma}$ is defined by $L_1 \cdot c L_2 = \bigcup_{t \in L_1} \{t\{c \leftarrow L_2\}\}$. The iterated c-product is inductively defined for $L \subset T_{\Sigma}$ by: $L^{0c} = \{c\}$ and $L^{(n+1)_c} = L^{n_c} \cup L \cdot c L^{n_c}$. The c-closure of L is defined by $L^{*c} = \bigcup_{n \geq 0} L^{n_c}$.

A regular expression over a ranked alphabet Σ is inductively defined by $E \in \Sigma_0, E = f(E_1, \dots, E_n), E = (E_1 + E_2), E = (E_1 \cdot_c E_2), E = (E_1^{*c}), where$ $c \in \Sigma_0, n \in \mathbb{N}, f \in \Sigma_n$ and E_1, E_2, \ldots, E_n are any n regular expressions over Σ . Parenthesis can be omitted when there is no ambiguity. We write $E_1 = E_2$ if E_1 and E_2 graphically coincide. We denote by RegExp (Σ) the set of all regular expressions over Σ . Every regular expression E can be seen as a tree over the ranked alphabet $\Sigma \cup \{+,\cdot_c,*_c\}$ with $c \in \Sigma_0$ where + and \cdot_c can be seen as a symbol of rank 2 and $*_c$ has rank 1. This tree is the syntax-tree $T_{\rm E}$ of E. The alphabetical width ||E|| of E is the number of occurrences of symbols of Σ in E. The size |E| of E is the size of its syntax tree T_E . The language [E] denoted by E is inductively defined as $\llbracket c \rrbracket = \{c\}, \llbracket f(E_1, E_2, \cdots, E_n) \rrbracket = f(\llbracket E_1 \rrbracket, \cdots, \llbracket E_n \rrbracket),$ $[[E_1 + E_2]] = [[E_1]] \cup [[E_2]], [[E_1 \cdot_c E_2]] = [[E_1]] \cdot_c [[E_2]], [[E_1^{*c}]] = [[E_1]]^{*c} \text{ where } n \in \mathbb{N},$ E_1, E_2, \ldots, E_n are any n regular expressions, $f \in \Sigma_n$ and $c \in \Sigma_0$. It is well known that a tree language is accepted by some tree automaton if and only if it can be denoted by a regular expression [5,8]. A regular expression E defined over Σ is linear if and only if every symbol of $\Sigma_{\geq 1}$ appears at most once in E. Note that any constant symbol may occur more than once. Let E be a regular expression over Σ . The linearized regular expression \overline{E}^{E} in E of a regular expression E is obtained from E by marking differently all symbols of a rank greater than or equal to 1 (symbols of $\Sigma_{\geq 1}$). The set of marked symbols with symbols of Σ_0 is the ranked alphabet containing symbols called positions. We denote this set by $Pos_{E}(E)$. When there is no ambiguity we denote by \overline{F} the subexpression \overline{F}^{E} with F is a subexpression of E. The mapping h is defined from $Pos_{E}(E)$ to Σ with $h(\operatorname{Pos}_{\mathrm{E}}(\mathrm{E})_m) \subset \Sigma_m$ for every $m \in \mathbb{N}$. It associates with a marked symbol $f_i \in$ Pos_E (E)_{>1} the symbol $f \in \Sigma_{\geq 1}$ and for a symbol $c \in \Sigma_0$ the symbol h(c) = c. We can extend the mapping h naturally to $\operatorname{RegExp}(\operatorname{Pos}_{E}(E)) \to \operatorname{RegExp}(\Sigma)$ by h(a) = a, $h(E_1 + E_2) = h(E_1) + h(E_2)$, $h(E_1 \cdot_c E_2) = h(E_1) \cdot_c h(E_2)$, $h(E_1^{*c}) = h(E_1) \cdot_c h(E_2)$ $h(\mathbf{E}_1)^{*_c}, h(f_j(\mathbf{E}_1, \dots, \mathbf{E}_n)) = f(h(\mathbf{E}_1), \dots, h(\mathbf{E}_n)), \text{ with } n \in \mathbb{N}, a \in \Sigma_0, f \in \Sigma_n,$ $f_j \in \text{Pos}_{\mathcal{E}}(\mathcal{E})_n$ such that $h(f_j) = f$ and $\mathcal{E}_1, \dots, \mathcal{E}_n$ any regular expressions over $Pos_{E}(E)$.

Example 1. Let $\Sigma_0 = \{a, c\}$, $\Sigma_1 = \{g, h\}$, $\Sigma_2 = \{f\}$ and $\Sigma = \Sigma_0 \cup \Sigma_1 \cup \Sigma_2$ be a ranked alphabet. Let E, F, G be the three following regular expressions over Σ : $F = ((c+a) + (g(c))^{*c})^{*c}$, G = f(a, h(c)) and $E = F \cdot_c G$. The linearized forms of E and G are: $\overline{E}^E = ((c+a) + (g_1(c))^{*c})^{*c} \cdot_c f_2(a, h_3(c))$, $\overline{G}^G = f_1(a, h_2(c))$.

The linearized form of G in E is $\overline{G}^{E} = f_2(a, h_3(c))$. Notice that Pos_G (G) = $\{a, f_1, h_2\} \neq \text{Pos}_{\mathbf{E}}(\mathbf{G}) = \{a, f_2, h_3\}$

3 Tree Automata Computations

In this section, we recall how to compute from a regular expression E a tree automaton that accepts [E]. We first recall the computation of the equation automaton \mathcal{A}_E of E, then we define the k-c-continuation automaton \mathcal{C}_E .

3.1 The Equation Tree Automaton

In [8], Kuske and Meinecke extend the notion of word partial derivatives [1] to tree partial derivatives in order to compute from a regular expression E a tree automaton recognizing [E]. Due to the notion of ranked alphabet, partial derivatives are no longer sets of expressions, but sets of tuples of expressions.

Let $\mathcal{N} = (E_1, \dots, E_n)$ be a tuple of regular expressions, F be some regular expression and $c \in \Sigma_0$. Then $\mathcal{N} \cdot_c F$ is the tuple $(E_1 \cdot_c F, \ldots, E_n \cdot_c F)$. For \mathcal{S} a set of tuples of regular expressions, $S \cdot_c F$ is the set $S \cdot_c F = \{ \mathcal{N} \cdot_c F \mid \mathcal{N} \in S \}$. Finally, $SET(\mathcal{N}) = \{E_1, \dots, E_m\}$ and $SET(\mathcal{S}) = \bigcup_{\mathcal{N} \in \mathcal{S}} SET(\mathcal{N})$.

Definition 1 ([8]). Let E be a regular expression over a ranked alphabet Σ and f be a symbol in Σ_m with $m \geq 1$ an integer. The set $f^{-1}(E)$ of tuples of regular expressions is defined as follows:

is is defined as follows:
$$f^{-1}(g(\mathbf{E}_{1},\cdots,\mathbf{E}_{n})) = \begin{cases} \{(\mathbf{E}_{1},\cdots,\mathbf{E}_{n})\} & \text{if } f = g \\ \emptyset & \text{otherwise} \end{cases}$$
$$f^{-1}(\mathbf{F}+\mathbf{G}) = f^{-1}(\mathbf{F}) \cup f^{-1}(\mathbf{G})$$
$$f^{-1}(\mathbf{F} \cdot_{c} \mathbf{G}) = \begin{cases} f^{-1}(\mathbf{F}) \cdot_{c} \mathbf{G} & \text{if } c \notin \llbracket \mathbf{F} \rrbracket \\ f^{-1}(\mathbf{F}) \cdot_{c} \mathbf{G} \cup f^{-1}(\mathbf{G}) & \text{otherwise} \end{cases}$$
$$f^{-1}(\mathbf{F}^{*c}) = f^{-1}(\mathbf{F}) \cdot_{c} \mathbf{F}^{*c}$$

The function f^{-1} is extended to any set S of regular expressions as follows: $f^{-1}(S) = \bigcup_{E \in S} f^{-1}(E).$

The partial derivative of E w.r.t. a word $w \in \Sigma_{>1}^*$, denoted by $\partial_w(E)$, is the set of regular expressions inductively defined by:

$$\partial_w(\mathbf{E}) = \begin{cases} \{\mathbf{E}\} & \text{if } w = \varepsilon \\ \mathrm{SET}(f^{-1}(\partial_u(\mathbf{E}))) & \text{if } w = uf, f \in \Sigma_{\geq 1}, u \in \Sigma_{\geq 1}^* \end{cases}$$
 The partial derivation is extended to any subset U of $\Sigma_{\geq 1}^*$ as by $\partial_U(\mathbf{E}) = U$

 $\bigcup_{w\in U} \partial_w(\mathbf{E}). \text{ Note that } \partial_{uf}(\mathbf{E}) = \partial_f(\partial_u(\mathbf{E})) = \bigcup_{\mathbf{F}\in\partial_u(\mathbf{E})} \partial_f(\widehat{\mathbf{F}}).$

Definition 2. Let E be a regular expression over a ranked alphabet Σ . The Equation Automaton of E is the tree automaton $A_{\rm E} = (Q, \Sigma, Q_T, \Delta)$ defined by

Equation Automaton of E is the tree automaton
$$\mathcal{A}_{E} = (Q, \mathcal{L}, Q_{T}, \Delta)$$
 defined
$$Q = \partial_{\Sigma_{\geq 1}^{*}}(E), \ Q_{T} = \{E\}, \ and$$

$$\Delta = \begin{cases} \{(F, f, G_{1}, \dots, G_{m}) \mid F \in Q, f \in \Sigma_{m}, m \geq 1, (G_{1}, \dots, G_{m}) \in f^{-1}(F)\} \\ \cup \{(F, c) \mid c \in (\llbracket F \rrbracket \cap \Sigma_{0})\} \end{cases}$$
Theorem 1. ([8]) Let E be a resolve conversion and A be the expection to

Theorem 1 ([8]). Let E be a regular expression and A_E be the equation tree automaton associated with E. Then $\mathcal{L}(\mathcal{A}_{E}) = [\![E]\!]$.

3.2 The C-Continuation Tree Automaton

In [8], Kuske and Meinecke show how to efficiently compute the equation tree automaton of a regular expression via an extension of Champarnaud and Ziadi's k-C-Continuation [3,4,7]. In this section, we show how to inductively compute them. The main difference with [8] is that the k-c-continuations are here computed using alternative formulae, and not using the partial derivation. As a consequence, any symbol that appears in the expression E admits a non-empty k-c-continuation (e.q. in [8], there is no continuation for g in $E = a \cdot_b g(c)$).

Definition 3. Let E be linear. Let k and m be two integers such that $1 \le k \le m$. Let f be in $(\Sigma_{\mathbb{E}} \cap \Sigma_m)$. The k-C-continuation $C_{f^k}(\mathbb{E})$ of f in E is the regular expression defined by:

$$C_{f^k}(g(\mathbf{E}_1, \dots, \mathbf{E}_m)) = \begin{cases} \mathbf{E}_k & \text{if } f = g \\ C_{f^k}(\mathbf{E}_j) & \text{if } f \in \Sigma_{\mathbf{E}_j} \end{cases}$$

$$C_{f^k}(\mathbf{F} + \mathbf{G}) = \begin{cases} C_{f^k}(\mathbf{F}) & \text{if } f \in \Sigma_{\mathbf{F}} \\ C_{f^k}(\mathbf{G}) & \text{if } f \in \Sigma_{\mathbf{G}} \end{cases}$$

$$C_{f^k}(\mathbf{F} \cdot_c \mathbf{G}) = \begin{cases} C_{f^k}(\mathbf{F}) \cdot_c \mathbf{G} & \text{if } f \in \Sigma_{\mathbf{F}} \\ C_{f^k}(\mathbf{G}) & \text{otherwise} \end{cases}$$

$$C_{f^k}(\mathbf{F}^{*_c}) = C_{f^k}(\mathbf{F}) \cdot_c \mathbf{F}^{*_c}$$

By convention, we set $C_{\varepsilon^1}(E) = E$.

Let us first show the relation between partial derivation and k-c-continuation.

Lemma 1. Let E be linear, n, m and k be three integers such that $n, m \geq 1$, $1 \leq k \leq m$, $f \in \Sigma_n$ and $g \in \Sigma_m \cup \{\varepsilon\}$. If $f^{-1}(C_{g^k}(E)) \neq \emptyset$ then $f^{-1}(C_{g^k}(E)) = \{(C_{f^1}(E), \ldots, C_{f^n}(E))\}$.

Proof. By induction over the structure of E. For any symbol $g \in \Sigma_p \cup \{\varepsilon\}$ and for any expression F, let us set $C_g(F) = (C_{g^1}(F), \dots, C_{g^p}(F))$.

- 1. Let us suppose that $E = h(E_1, \dots, E_m)$. Three cases have to be considered:
 - (a) If $g = \varepsilon$, then k = 1 and $f^{-1}(C_{g^k}(E)) = f^{-1}(E)$. Since $f^{-1}(C_{g^k}(E)) \neq \emptyset$, f = h. Hence, $f^{-1}(E) = \{(E_1, \dots, E_n)\}$. Moreover, for any integer $1 \leq j \leq n$, $C_{f^j}(E) = E_j$. Consequently, $f^{-1}(C_{g^k}(E)) = \{C_f(E)\}$.
 - (b) Let us suppose that $g \neq \varepsilon$ and $g \neq h$. Hence $C_{g^k}(\mathbf{E}) = C_{g^k}(\mathbf{E}_l)$ with $f \in \Sigma_{\mathbf{E}_l}$. By induction hypothesis, $f^{-1}(C_{g^k}(\mathbf{E}_j)) = \{C_f(\mathbf{E}_l)\}$. Moreover, for any integer $1 \leq j \leq n$, $C_{f^j}(\mathbf{E}) = C_{f^j}(\mathbf{E}_l)$. Consequently, $f^{-1}(C_{g^k}(\mathbf{E})) = \{C_f(\mathbf{E})\}$.
 - (c) Let us suppose that $g \neq \varepsilon$ and g = h. Hence $C_{g^k}(\mathbf{E}) = \mathbf{E}_k$. Since $f^{-1}(C_{g^k}(\mathbf{E})) \neq \emptyset$, then $f \in \Sigma_{E_k}$. Thus, $f \neq h$. By definition, $\mathbf{E}_k = C_{\varepsilon}^1(\mathbf{E}_k)$. By induction hypothesis, $f^{-1}(C_{\varepsilon}^1(\mathbf{E}_k)) = \{C_f(\mathbf{E}_k)\}$. Since $f \neq h$ and since $f \in \Sigma_{E_k}$, for any integer $1 \leq j \leq n$, $C_{f^j}(\mathbf{E}) = C_{f^j}(\mathbf{E}_k)$. Consequently, $f^{-1}(C_{g^k}(\mathbf{E})) = \{C_f(\mathbf{E})\}$.
- 2. Suppose that $E = E_1 + E_2$. Suppose that $f \in \Sigma_{E_1}$. Then $f^{-1}(C_{g^k}(E)) = f^{-1}(C_{g^k}(E_1))$. By induction hypothesis, $f^{-1}(C_{g^k}(E_1)) = \{C_f(E_1)\}$. Finally, since for any integer $1 \le j \le n$, $C_{f^j}(E) = C_{f^j}(E_1)$, it holds $f^{-1}(C_{g^k}(E)) = \{C_f(E)\}$. The prove is identical whenever $f \in \Sigma_{E_2}$.

- 3. Let us suppose that $E = E_1 \cdot_c E_2$. Two cases have to be considered:
 - (a) If $g \in \Sigma_{E_1}$, $f^{-1}(C_{g^k}(E)) = f^{-1}(C_{g^k}(E_1) \cdot_c E_2)$. If $f \in \Sigma_{E_1}$, then $f^{-1}(C_{g^k}(E_1) \cdot_c E_2) = f^{-1}(C_{g^k}(E_1) \cdot_c E_2) = f^{-1}(E_2)$. Hence, according to induction hypothesis, either $f^{-1}(C_{g^k}(E_1) \cdot_c E_2) = \{(C_{f^1}(E_1) \cdot_c E_2, \dots, C_{f^n}(E)) \cdot_c E_2\}$, or $f^{-1}(C_{g^k}(E_1) \cdot_c E_2) = \{(C_{f^1}(E_2), \dots, C_{f^n}(E_2))\}$. By definition, considering whether $f \in \Sigma_{E_1}$, for any integer $1 \leq j \leq n$, either $C_{f^j}(E) = C_{f^1}(E_1) \cdot_c E_2$ or $C_{f^j}(E) = C_{f^1}(E_2)$. In both of these cases, $f^{-1}(C_{g^k}(E)) = \{C_f(E)\}$.
 - (b) If $g \in \Sigma_{E_2}$, $f^{-1}(C_{g^k}(E)) = f^{-1}(C_{g^k}(E_2)$. By induction hypothesis, $f^{-1}(C_{g^k}(E_2)) = \{(C_{f^1}(E_2), \dots, C_{f^n}(E_2))\}$. Moreover, for any integer $1 \le j \le n$, $C_{f^j}(E) = C_{f^j}(E_2)$. Consequently, $f^{-1}(C_{g^k}(E)) = \{C_f(E)\}$.
- 4. Let us suppose that $E = E_1^{*c}$. Two cases have to be considered:
 - (a) If $g = \varepsilon$, then $f^{-1}(C_{g^k}(E)) = f^{-1}(E_1^{*c}) = f^{-1}(E_1) \cdot_c E_1^{*c}$. By definition, $E_1 = C_{\varepsilon}(E_1)$. Hence by induction hypothesis, $f^{-1}(C_{\varepsilon}(E_1)) \cdot_c E_1^{*c} = \{(C_{f^1}(E_1) \cdot_c E_1^{*c}, \dots, C_{f^n}(E_1) \cdot_c E_1^{*c}\}$. Moreover, for any integer $1 \le j \le n$, $C_{f^j}(E) = C_{f^j}(E_1) \cdot_c E_1^{*c}$. Consequently, $f^{-1}(C_{g^k}(E)) = \{C_f(E)\}$.
 - (b) Suppose that $g \neq \varepsilon$. Then $f^{-1}(C_{g^k}(E)) = f^{-1}(C_{g^k}(E_1) \cdot_c E_{1^c}^{*c})$. Depending whether c belongs to $[\![C_{g^k}(E_1)]\!]$, either $f^{-1}(C_{g^k}(E_1) \cdot_c E_{1^c}^{*c}) = f^{-1}(C_{g^k}(E_1)) \cdot_c E_{1^c}^{*c}$ or $f^{-1}(C_{g^k}(E_1) \cdot_c E_{1^c}^{*c}) = f^{-1}(C_{g^k}(E_1)) \cdot_c E_{1^c}^{*c} \cup f^{-1}(E_{1^c}^{*c})$. Since $E_1^{*c} = C_{\varepsilon}^1(E_{1^c}^{*c})$, it holds by induction hypothesis that either $f^{-1}(C_{g^k}(E_1) \cdot_c E_{1^c}^{*c}) = \{(C_{f^1}(E_1) \cdot_c E_{1^c}^{*c}, \dots, C_{f^n}(E_1) \cdot_c E_{1^c}^{*c})\}$ or $f^{-1}(C_{g^k}(E_1) \cdot_c E_{1^c}^{*c}) = \{(C_{f^1}(E_1) \cdot_c E_{1^c}^{*c}, \dots, C_{f^n}(E_1) \cdot_c E_{1^c}^{*c})\}$. Finally, since for any integer $1 \leq j \leq n$, $C_{f^j}(E) = C_{f^j}(E_1) \cdot_c E_{1^c}^{*c}$, in both of these cases, $f^{-1}(C_{g^k}(E)) = \{C_f(E)\}$.

Proposition 1. Let E be linear and $f \in \Sigma_n$ with $n \ge 1$. Let u be a word in $\Sigma_{\ge 1}^*$. If $f^{-1}(\partial_u(E)) \ne \emptyset$ then $f^{-1}(\partial_u(E)) = \{(C_{f^1}(E), \ldots, C_{f^n}(E))\}.$

Proof. By recurrence over the length of u. For any symbol $g \in \Sigma_p \cup \{\varepsilon\}$ and for any expression F, let us set $C_q(F) = (C_{q^1}(F), \dots, C_{q^p}(F))$.

- 1. Let $u = \varepsilon$. Then $f^{-1}(\partial_u(\mathbf{E})) = f^{-1}(\mathbf{E})$. By definition, $f^{-1}(\mathbf{E}) = f^{-1}(C_{\varepsilon}^1(\mathbf{E}))$. According to Lemma 1, $f^{-1}(C_{\varepsilon}^1(\mathbf{E})) = \{C_f(\mathbf{E})\}$.
- 2. Let u=wg with w a word in $\Sigma_{\geq 1}^*$ and g a symbol in Σ_m . Then $f^{-1}(\partial_u(\mathbf{E}))=f^{-1}(SET(g^{-1}(\partial_u(\mathbf{E})))$. According to recurrence hypothesis, it holds that $SET(g^{-1}(\partial_u(\mathbf{E})))=SET(\{C_g(\mathbf{E})\})=\{C_{g^1}(\mathbf{E}),\ldots,C_{g^m}(\mathbf{E})\}$. By definition, $f^{-1}(\{C_{g^1}(\mathbf{E}),\ldots,C_{g^m}(\mathbf{E})\})=\bigcup_{1\leq i\leq m}f^{-1}(C_{g^i}(\mathbf{E}))$. According to Lemma 1, for any integer i such that $f^{-1}(C_{g^i}(\mathbf{E}))\neq\emptyset$, it holds $f^{-1}(C_{g^i}(\mathbf{E}))=\{(C_{f^1}(\mathbf{E}),\ldots,C_{f^n}(\mathbf{E}))\}$. Since $f^{-1}(\partial_u(\mathbf{E}))\neq\emptyset$, there exists at least one integer i such that $f^{-1}(C_{g^i}(\mathbf{E}))\neq\emptyset$. Consequently, $\bigcup_{1\leq i\leq m}f^{-1}(C_{g^i}(\mathbf{E}))=\{C_f(\mathbf{E})\}$.

Definition 4. The automaton $\overline{\mathcal{C}}_{\mathrm{E}} = (Q_{\overline{\mathcal{C}}}, \operatorname{Pos}_{\mathrm{E}}(\mathrm{E}), \{C_{\varepsilon^1}(\overline{\mathrm{E}})\}, \Delta_{\overline{\mathcal{C}}})$ is defined by $-Q_{\overline{\mathcal{C}}} = \{C_{f_j^k}(\overline{\mathrm{E}}) \mid f_j \in \operatorname{Pos}_{\mathrm{E}}(\mathrm{E})_m, 1 \leq k \leq m\} \cup \{C_{\varepsilon^1}(\overline{\mathrm{E}})\},$

$$- \Delta_{\overline{C}} = \begin{cases} \{ (C_x(\overline{\mathbf{E}}), g_i, \mathfrak{C}_{g_i}) \mid g_i \in \operatorname{Pos_E}(\mathbf{E})_m, m \ge 1, \mathfrak{C}_{g_i} \in g_i^{-1}(C_x(\overline{\mathbf{E}})) \} \\ \cup \{ (C_x(\overline{\mathbf{E}}), c) \mid, c \in \llbracket C_x(\overline{\mathbf{E}}) \rrbracket \cap \Sigma_0 \} \end{cases}$$

where for any symbol g_i in $\operatorname{Pos}_{\mathbf{E}}(\mathbf{E})_m$, $\mathfrak{C}_{g_i} = (C_{g_i^1}(\overline{\mathbf{E}}), \dots, C_{g_i^m}(\overline{\mathbf{E}}))$.

The following lemma illustrates the link between $\overline{\mathcal{C}}_{E}$ and $\mathcal{A}_{\overline{E}}$:

Lemma 2. The coaccessible part of $\overline{\mathcal{C}}_{E}$ is equal to $\mathcal{A}_{\overline{E}}$.

Proof. The expression $\overline{\mathbb{E}}$ is the final state of the two automata. Let us suppose now that q is a coaccessible state both in $\overline{\mathcal{C}}_{\mathrm{E}}$ and $\mathcal{A}_{\overline{\mathrm{E}}}$. Hence, from Definition 4 and from Definition 2:

there exists a transition (q, f, q_1, \ldots, q_n) in $\mathcal{A}_{\overline{\mathbb{R}}}$

- $\Leftrightarrow (q_1,\ldots,q_n)\in f^{-1}(q)$
- $\Leftrightarrow (q_1,\ldots,q_n)=(C_f^1(\overline{E}),\ldots,C_f^n(\overline{E}))\in f^{-1}(q)$ (Proposition 1)
- \Leftrightarrow there exists a transition (q, f, q_1, \ldots, q_n) in $\overline{\mathcal{C}}_{\mathrm{E}}$.

Hence, the states q_1, \ldots, q_n are coaccessible from q by f in $\overline{\mathcal{C}}_E$ if and only if they are in $\mathcal{A}_{\overline{E}}$. Consequently, the coaccessible part of $\overline{\mathcal{C}}_E$ is equal to the equation tree automaton $\mathcal{A}_{\overline{E}}$.

Corollary 1. The automaton $\overline{\mathcal{C}}_{\mathrm{E}}$ accepts $[\![\overline{\mathrm{E}}]\!]$.

The *C-Continuation tree automaton* $C_{\mathbb{E}}$ associated with \mathbb{E} is obtained by replacing each transition $(C_x(\overline{\mathbb{E}}), g_i, C_{g_i^1}(\overline{\mathbb{E}}), \dots, C_{g_i^m}(\overline{\mathbb{E}}))$ of the tree automaton $\overline{\mathcal{C}}_{\mathbb{E}}$ by $(C_x(\overline{\mathbb{E}}), h(g_i), C_{g_i^1}(\overline{\mathbb{E}}), \dots, C_{g_i^m}(\overline{\mathbb{E}}))$.

Corollary 2.
$$h(\mathcal{L}(\overline{\mathcal{C}}_{\mathrm{E}})) = \mathcal{L}(\mathcal{C}_{\mathrm{E}}) = [\![\mathrm{E}]\!].$$

In what follows, for any two trees s and t, we denote by $s \leq t$ the relation "s is a subtree of t". Let k be an integer. We denote by root(s) the root of any tree s and by k-child(t), for a tree $t = f(t_1, \ldots, t_n)$, the k^{th} child of f in t that is root of t_k if it exists.

Let $1 \leq k \leq m$ be two integers and f_j be a symbol in $\operatorname{Pos_E}(\mathbf{E})_m$. The sets $\operatorname{First}(\mathbf{E})$ is the subset of $\operatorname{Pos_E}(\mathbf{E})$ defined by $\operatorname{First}(\mathbf{E}) = \{\operatorname{root}(t) \in \operatorname{Pos_E}(\mathbf{E}) \mid t \in [\overline{\mathbf{E}}]\}$. The set $\operatorname{Follow}(\mathbf{E}, f_j, k)$ is the subset of $\operatorname{Pos_E}(\mathbf{E})$ defined by $\operatorname{Follow}(\mathbf{E}, f_j, k) = \{g_i \in \operatorname{Pos_E}(\mathbf{E}) \mid \exists t \in [\overline{\mathbf{E}}]\}, \exists s \leq t, \operatorname{root}(s) = f_j, k\text{-child}(s) = g_i\}$.

Proposition 2 ([9]). The computation of all the sets $(Follow(E, f_j, k))_{1 \le k \le m, f \in Pos_E(E)_m}$ can be done with an O(|E|) time and space complexity.

Proposition 3. Let $1 \le k \le m$ be two integers and f_j be a position in $\operatorname{Pos_E}(E)_m$. If $\operatorname{Follow}(E, f_j, k) \ne \emptyset$ then $\operatorname{Follow}(E, f_j, k) = \operatorname{First}(C_{f_j^k}(\overline{E}))$.

Proof. Let E be a linear regular expression over a ranked alphabet Σ , $1 \le k \le m$ be two integers and f be a symbol in Σ_m . The set $\lambda^f(E, k)$ is the subset of Σ_0 defined by $\lambda^f(E, k) = \{c \in \Sigma_0 \mid \exists t \in \llbracket E \rrbracket, \exists f(t_1, \ldots, t_m) \le t, t_k = c\}$. The set $\lambda(E)$ is the subset of Σ_0 defined by $\lambda(E) = \bigcup_{g \in \Sigma_m, 1 \le k \le m} \lambda^g(E, k)$.

 $\lambda(E)$ is the subset of Σ_0 defined by $\lambda(E) = \bigcup_{g \in \Sigma_m, 1 \le k \le m} \lambda^g(E, k)$. Let E be a regular expression over a ranked alphabet Σ , $1 \le k \le m$ be two integers and f_i be a symbol in $Pos_E(E)_m$. In [9], it is shown, using alternative and equivalent formulae, that the set $\operatorname{Follow}(E, f_j, k)$ is equal to $\operatorname{Follow}(\overline{E}, f_j, k)$, where $\operatorname{Follow}(F, f_j, k)$ is the subset of $\operatorname{Pos}_E(E)$ inductively defined for any linear regular expression F as follows:

regular expression
$$F$$
 as follows: Follow $(a, f, k) = \emptyset$,

Follow $(E_1 + E_2, f, k) = \begin{cases} \text{Follow}(E_1, f, k) & \text{if } f \in \Sigma_{E_1}, \\ \text{Follow}(E_2, f, k) & \text{if } f \in \Sigma_{E_2}, \end{cases}$

Follow $(E_1 \cdot c E_2, f, k) = \begin{cases} (\text{Follow}(E_1, f, k) \setminus \{c\}) \cup \text{First}(E_2) & \text{if } c \in \lambda^f(E_1, k), \\ \text{Follow}(E_1, f, k) & \text{if } f \in \Sigma_{E_1} \land c \notin \lambda^f(E_1, k), \\ \text{Follow}(E_2, f, k) & \text{if } f \in \Sigma_{E_2} \land c \in \lambda(E_1), \\ \emptyset & \text{otherwise}, \end{cases}$

Follow $(E_1^*, f, k) = \begin{cases} \text{Follow}(E_1, f, k) \cup \text{First}(E_1) & \text{if } c \in \lambda(E_1), \\ \text{Follow}(E_1, f, k) & \text{otherwise}, \end{cases}$

Follow $(g(E_1, \dots, E_n), f, k) = \begin{cases} \text{First}(E_k) & \text{if } f = g, \\ \text{Follow}(E_l, f, k) & \text{if } f \in \Sigma_{E_l}. \end{cases}$

Since by definition $\operatorname{Follow}(E, f_j, k) = \operatorname{Follow}(\overline{E}, f_j, k)$, let us show by induction over \overline{E} that if $\operatorname{Follow}(E, f_j, k) \neq \emptyset$ then $\operatorname{Follow}(E, f_j, k) = \operatorname{First}(C_{f_j^k}(\overline{E}))$. Let us set $\overline{E} = F$.

Suppose that $F = f_j(F_1, \ldots, F_m)$. Hence $\operatorname{Follow}(F, f_j, k) = \operatorname{First}(F_k)$. Moreover by definition $C_{f_j^k}(F) = F_k$. Then $\operatorname{Follow}(F, f_j, k) = \operatorname{First}(C_{f_j^k}(F))$. The property is true for the base case.

Assuming that the property holds for the subexpressions of F.

- 1. Consider that $F = g_i(F_1, \ldots, F_m)$ with $f_j \neq g_i$. Then by definition $\operatorname{Follow}(F, f_j, k) = \operatorname{Follow}(F_l, f_j, k)$ with $f_j \in \Sigma_{F_l}$. By induction hypothesis, $\operatorname{Follow}(F_l, f_j, k) = \operatorname{First}(C_{f_j^k}(F_l))$. Moreover, from Definition 3, $C_{f_j^k}(F) = C_{f_j^k}(F_l)$. Consequently, $\operatorname{Follow}(F, f_j, k) = \operatorname{First}(C_{f_j^k}(F))$.
- 2. Let us consider that $F = F_1 + F_2$. Suppose that $f_j \in \Sigma_{F_i}$ with $i \in \{1, 2\}$. Hence $\operatorname{Follow}(F_1 + F_2, f_j, k) = \operatorname{Follow}(F_i, f_j, k)$. By induction hypothesis, $\operatorname{Follow}(F_i, f_j, k) = \operatorname{First}(C_{f_j^k}(F_i))$. From Definition 3, $C_{f_j^k}(F_1 + F_2) = C_{f_j^k}(E_i)$. Consequently, $\operatorname{Follow}(F_1 + F_2, f_j, k) = \operatorname{First}(C_{f_i^k}(F_1 + F_2))$.
- 3. Consider that $F = F_1 \cdot_c F_2$. Three cases may occur.
 - (a) Suppose that $c \in \lambda^{f_j}(F_1, k)$. Then Follow $(F_1 \cdot_c F_2, f_j, k) = (\text{Follow}(F_1, f_j, k) \setminus \{c\}) \cup \text{First}(F_2)$. By induction hypothesis, Follow $(F_1 \cdot_c F_2, f_j, k) = (\text{First}(C_{f_j^k}(F_1)) \setminus \{c\}) \cup \text{First}(F_2)$. Moreover, from Definition 3, $C_{f_j^k}(F) = C_{f_j^k}(F_1) \cdot_c F_2$. Since $c \in \lambda^{f_j}(F_1, k)$, then by definition of $\lambda^{f_j}(F_1, k)$, $c \in \text{Follow}(F_1, f_j, k)$. By induction hypothesis, Follow $(F_1, f_j, k) = \text{First}(C_{f_j^k}(F_1))$ and then $c \in [C_{f_j^k}(F_1)]$. Consequently, by definition, $\text{First}(C_{f_j^k}(F_1) \cdot_c F_2) = (\text{First}(C_{f_j^k}(F_1)) \setminus \{c\}) \cup \text{First}(F_2)$. Therefore, Follow $(F_1 \cdot_c F_2, f_j, k) = \text{First}(C_{f_j^k}(F_1 \cdot_c F_2))$.
 - (b) Consider that $c \notin \lambda^{f_j}(F_1, k)$ and $f_j \in \Sigma_{F_1}$. In this case Follow $(F_1 \cdot c F_2, f_j, k) = \text{Follow}(F_1, f_j, k)$. By induction hypothesis, Follow $(F_1, f_j, k) = \text{First}(C_{f_j^k}(F_1))$. From Definition 3, $C_{f_j^k}(F_1 \cdot c F_2) = C_{f_j^k}(F_1) \cdot c F_2$. Since $c \notin \lambda^{f_j}(F_1, k)$, then by definition $c \notin \text{Follow}(F_1, f_j, k)$. By induction hypothesis, Follow $(F_1, f_j, k) = \text{First}(C_{f_j^k}(F_1))$ and then $c \notin [C_{f_j^k}(F_1)]$.

- Consequently, $\operatorname{First}(C_{f_j^k}(F_1 \cdot_c F_2)) = \operatorname{First}(C_{f_j^k}(F_1))$. Then $\operatorname{Follow}(F_1 \cdot_c F_2, f_j, k) = \operatorname{First}(C_{f_j^k}(F_1 \cdot_c F_2))$.
- (c) Consider that $c \in \lambda(F_1)$ and $f_j \in \Sigma_{F_2}$. In this case Follow $(F_1 \cdot_c F_2, f_j, k) = \text{Follow}(F_2, f_j, k)$. By induction hypothesis, Follow $(F_2, f_j, k) = \text{First}(C_{f_j^k}(F_2))$. From Definition 3, $C_{f_j^k}(F_1 \cdot_c F_2) = C_{f_j^k}(F_2)$. Therefore, Follow $(F_1 \cdot_c F_2, f_j, k) = \text{First}(C_{f_j^k}(F_1 \cdot_c F_2))$.
- 4. Consider that $F = F_1^{*c}$. By Definition 3, $C_{f_j^k}(F_1^{*c}) = C_{f_j^k}(F_1) \cdot_c F_1$. Two cases may occur.
 - (a) Suppose that $c \in \lambda(F_1)$. In this case, Follow $(F_1^{*c}, f_j, k) = \text{Follow}(F_1, f_j, k) \cup \text{First}(F_1^{*c})$. By induction hypothesis, Follow $(F_1, f_j, k) = \text{First}(C_{f_j^k}(F_1))$. By definition, $c \in \text{Follow}(F_1, f_j, k)$. Since by induction Follow $(F_1, f_j, k) = \text{First}(C_{f_j^k}(F_1))$, $c \in \text{First}(C_{f_j^k}(F_1))$ and then $c \in [C_{f_j^k}(F_1)]$. Consequently, First $(C_{f_j^k}(F_1) \cdot c F_1) = \text{First}(C_{f_j^k}(F_1)) \cup \text{First}(F_1)$. Consequently Follow $(F_1^{*c}, f_j, k) = \text{Follow}(F_1, f_j, k) \cup \text{First}(F_1^{*c}) = \text{First}(C_{f_j^k}(F_1)) \cup \text{First}(F_1) = \text{First}(C_{f_j^k}(F_1^{*c}))$.
 - (b) Suppose that $c \notin \lambda(\tilde{F}_1)$. In this case, $\operatorname{Follow}(F_1^{*c}, f_j, k) = \operatorname{Follow}(F_1, f_j, k)$. By induction hypothesis, $\operatorname{Follow}(F_1, f_j, k) = \operatorname{First}(C_{f_j^k}(F_1))$. By definition, $c \notin \operatorname{Follow}(F_1, f_j, k)$. Since by induction $\operatorname{Follow}(F_1, f_j, k) = \operatorname{First}(C_{f_j^k}(F_1))$, $c \notin \operatorname{First}(C_{f_j^k}(F_1))$ and then $c \notin \llbracket C_{f_j^k}(F_1) \rrbracket$. Consequently, $\operatorname{First}(C_{f_j^k}(F_1) \cdot_c F_1) = \operatorname{First}(C_{f_j^k}(F_1))$. Consequently $\operatorname{Follow}(F_1^{*c}, f_j, k) = \operatorname{Follow}(F_1, f_j, k) = \operatorname{First}(C_{f_j^k}(F_1)) = \operatorname{First}(C_{f_j^k}(F_1^{*c}))$.

Proposition 4. Let $1 \le k \le m$ be two integers, f_j be a symbol in $\operatorname{Pos_E}(E)_m$ and g_i be a symbol in $\operatorname{Pos_E}(E)$. Then $g_i^{-1}(C_{f_i^k}(\overline{E})) \ne \emptyset \Leftrightarrow g_i \in \operatorname{First}(C_{f_i^k}(\overline{E}))$.

Proof. Let F be a linear expression. Let us show by induction over the structure of F that $g_i^{-1}(F) \neq \emptyset \Leftrightarrow g_i \in \text{First}(F)$.

- 1. Consider that $F = g_i(F_1, ..., F_n)$. By definition, First $(F) = \{g_i\}$. By definition, $g_i^{-1}(F) = \{(F_1, ..., F_n)\}$. Hence the two conditions are both satisfied.
- 2. Consider that $F = f(F_1, ..., F_n)$ with $f \in \Sigma_F \setminus \{g_i\}$. By definition, First $(F) = \{f\}$. By definition, $g_i^{-1}(F) = \emptyset$. Hence the two conditions are both unsatisfied.
- 3. If $F = F_1 + F_2$, then according to [9], $\operatorname{First}(F) = \operatorname{First}(F_1) \cup \operatorname{First}(F_2)$. By definition, $g_i^{-1}(F) = g_i^{-1}(F_1) \cup g_i^{-1}(F_2)$. By induction hypothesis, for $l \in \{1, 2\}, \ g_i^{-1}(F_l) \neq \emptyset \Leftrightarrow g_i \in \operatorname{First}(F_l)$. Consequently, $g_i^{-1}(F) \neq \emptyset \Leftrightarrow g_i \in \operatorname{First}(F_1) \cup \operatorname{First}(F_2)$.
- 4. If $F = F_1 \cdot_c F_2$, then according to [9], First $(F) = \text{First}(F_1) \cup (\text{First}(F_2) \mid c \in c \in \llbracket F_1 \rrbracket)$. By definition, $g_i^{-1}(F) = g_i^{-1}(F_1) \cdot_c F_2 \cup (g_i^{-1}(F_2) \mid c \in \llbracket F_1 \rrbracket)$. By induction hypothesis, for $l \in \{1, 2\}, g_i^{-1}(F_l) \neq \emptyset \Leftrightarrow g_i \in \text{First}(F_l)$. Consequently, $g_i^{-1}(F) \neq \emptyset \Leftrightarrow g_i \in \text{First}(F)$.
- 5. If $F = F_1^{*c}$, then according to [9], $\operatorname{First}(F) = \operatorname{First}(F_1)$. By definition, $g_i^{-1}(F) = g_i^{-1}(F_1) \cdot_c F_1^{*c}$. By induction hypothesis, $g_i^{-1}(F_1) \neq \emptyset \Leftrightarrow g_i \in \operatorname{First}(F_1)$. Consequently, $g_i^{-1}(F) \neq \emptyset \Leftrightarrow g_i \in \operatorname{First}(F)$.

As a direct consequence, the conditions of Proposition 4 are equivalent.

Lemma 3. Let $1 \leq k \leq m$ be two integers and f_j be a position in $\operatorname{Pos_E}(E)_m$. If $\operatorname{Follow}(E, f_j, k) = \emptyset$ then $C_{f_j^k}(\overline{E})$ is not a coaccessible state in C_E .

Proof. Let us first show that for any state $q=C_{f_j^k}(\overline{\mathbb{E}})$, there exists a tree t such that $q\in \Delta^*(t)$, where Δ is the transition function of $\overline{\mathcal{C}}_{\mathbb{E}}$ (proposition \mathbf{P} in the following). By definition, $[\![q]\!]$ is not empty. If there exists a constant $c\in [\![q]\!]$, then by construction $q\in \Delta^*(c)$. If $t=g_i(t_1,\ldots,t_n)\in [\![q]\!]$, then by definition $g_i\in \mathrm{First}(q)$. According to Proposition $4,g_i^{-1}(C_{f_j^k}(\overline{\mathbb{E}}))\neq\emptyset$. Furthermore, according to Lemma $1,g_i^{-1}(C_{f_j^k}(\overline{\mathbb{E}}))=\{(C_{g_i^1}(\overline{\mathbb{E}}),\ldots,C_{g_i^n}(\overline{\mathbb{E}}))\}$. Hence, the states $q_1=C_{g_i^1}(\overline{\mathbb{E}}),\ldots,q_n=C_{g_i^n}(\overline{\mathbb{E}})$ are coaccessible from q. By induction hypothesis, there exists a tree t_l' in $[\![q_l]\!]$ such that $q_l\in \Delta^*(t_l')$. As a direct consequence, $q\in \Delta^*(g_i(t_1',\ldots,t_n'))$.

Let us show that if q is coaccessible, then there exists a tree t in $\mathcal{L}(\overline{\mathcal{C}}_{\mathrm{E}})$ for any tree s satisfying $q \in \Delta^*(s)$ such that $s \preccurlyeq t$ (proposition \mathbf{P} ' in the following). If $q = C_{\varepsilon}^1(\overline{\mathbf{E}})$, any tree s such that $s \in \Delta^*(t)$ is accepted since q is final. Setting t = s, property holds. Otherwise, q is coaccessible from a state p. By construction, there exists a transition $(p, f_j, (q_1, \ldots, q_m))$ with $q_k = q$. By induction hypothesis, there exists a tree t in $\mathcal{L}(\overline{\mathcal{C}}_{\mathrm{E}})$ for any tree s' satisfying $p \in \Delta^*(s')$ such that $s' \preccurlyeq t$. Since any tree s satisfying $q \in \Delta^*(s)$, is a subtree of a tree s' satisfying $p \in \Delta^*(s')$ which root is f_j , there exists a tree t in $\mathcal{L}(\overline{\mathcal{C}}_{\mathrm{E}})$ for any tree s satisfying $q \in \Delta^*(s)$ such that $s \preccurlyeq t$.

Suppose that $q = C_{f_j^k}(\overline{\mathbf{E}})$ is a coaccessible state in $\overline{\mathcal{C}}_{\mathbf{E}}$. According to $(\mathbf{P'})$, there exists a tree t in $\mathcal{L}(\overline{\mathcal{C}}_{\mathbf{E}})$ for any tree s satisfying $q \in \Delta^*(s)$ such that $s \preccurlyeq t$. By construction, since $C_{f_j^k}(\overline{\mathbf{E}})$ is coaccessible by the symbol f_j , there exists a tree $s' \preccurlyeq t$ such that $\mathrm{root}(s') = f_j$ and $k - \mathrm{child}(s') = \mathrm{root}(s)$. By definition $\mathrm{root}(s) \in \mathrm{Follow}(\overline{\mathbf{E}}, f_j, k)$. As a direct consequence, if $C_{f_j^k}(\overline{\mathbf{E}})$ is a coaccessible state in $\mathcal{C}_{\mathbf{E}}$, it is in $\overline{\mathcal{C}}_{\mathbf{E}}$. As previously shown, this implies that $\mathrm{Follow}(\overline{\mathbf{E}}, f_j, k) \neq \emptyset$. As a conclusion, by definition, $\mathrm{Follow}(\overline{\mathbf{E}}, f_j, k) = \mathrm{Follow}(\mathbf{E}, f_j, k) \neq \emptyset$.

3.3 From k-C-Continuation Automaton to Equation Automaton

The equation automaton is a quotient of the C-Continuation one w.r.t. the equivalence relation denoted by \sim_e over the set of states of $\overline{\mathcal{C}}_{\mathrm{E}}$ defined for any two states $q_1 = C_{f_i^k}(\mathrm{E})$ and $q_2 = C_{g_i^p}(\mathrm{E})$ by $q_1 \sim_e q_2 \Leftrightarrow h(q_1) = h(q_2)$.

Proposition 5. The coaccessible part of the finite tree automaton $C_{\rm E}/_{\sim_e}$ is isomorphic to the equation tree automaton $A_{\rm E}$.

Proof. Let E be a regular expression over an alphabet Σ . We define the inverse function of h denoted by $h^{-1}: \Sigma \to 2^{\operatorname{Pos_E}(E)}$ such that for any symbol x in Σ , $h^{-1}(x) = \{y \in \operatorname{Pos_E}(E) \mid h(y) = x\}$.

Theorem 2 ([8]). Let E be a regular expression over an alphabet Σ . Then for every $u \in \Sigma_{\geq 1}^*$, $\partial_u(E) = \bigcup_{\overline{u} \in h^{-1}(u)} h(\partial_{\overline{u}}(\overline{E}))$.

Proposition 6 ([8]). Let E be a regular expression over a ranked alphabet Σ . Then we have for every $f \in \Sigma_{\geq 1}$,

$$f^{-1}(\mathbf{E}) = \bigcup_{f_j \in h^{-1}(f)} h(f_j^{-1}(\overline{\mathbf{E}})) \text{ and } \partial_f(\mathbf{E}) = \bigcup_{f_j \in h^{-1}(f)} h(\partial_{f_j}(\overline{\mathbf{E}}))$$

Let us denote by $\mathfrak C$ the coaccessible part of the finite tree automaton $C_{\mathbb E}/_{\sim_e}$. Let Q be the set of states of $\mathcal A_{\mathbb E}$ and Q' be the set of states of $\mathfrak C$. The isomorphism of the sets of states can be shown by the function $\phi:Q'\to Q:$ $[C_{f_j^k}(\mathbb E)]\mapsto h(C_{f_j^k}(\mathbb E))$. Indeed, according to Lemma $2,C_{f_j^k}(\mathbb E)\in\partial_{\overline u}(\overline{\mathbb E})$ for some $\overline u\in\operatorname{Pos}_{\mathbb E}(\mathbb E)^*_{\geq 1}$. Using Theorem $2,h(C_{f_j^k}(\mathbb E))=h(\partial_{\overline u}(\overline{\mathbb E}))\in\partial_{\Sigma^*_{\geq 1}}(\mathbb E)=Q.$ Injectivity of ϕ can be shown directly from the definition of the equivalence relation \sim_e . For surjectivity, it is deduced from the Theorem 2.

By definition, $\phi([C_{\varepsilon^1}(E)]) = C_{\varepsilon^1}(E) = E$. Hence the image of the final state of \mathfrak{C} is the final state of \mathcal{A}_E .

Let us show that the sets of transitions are also isomorphic.

Let $([C_{f_j}^k(\mathbf{E})], g, [C_{g_i^1}(\mathbf{E})] \dots, [C_{g_i^m}(\mathbf{E})])$ be a transition in \mathfrak{C} . Equivalently by construction, there exists a symbol g_i such that $(C_{f_j}^k(\mathbf{E}), g_i, C_{g_i^1}(\mathbf{E}) \dots, C_{g_i^m}(\mathbf{E}))$ is a transition in the accessible part of the automaton $\overline{\mathcal{C}}_{\mathbf{E}}$. As the coaccessible part of $\overline{\mathcal{C}}_{\mathbf{E}}$ and $\mathcal{A}_{\overline{\mathbf{E}}}$ are equal (by Lemma 2), the transition $(\overline{\mathbf{F}}, g_i, \overline{\mathbf{H}}_1, \dots, \overline{\mathbf{H}}_m)$ is in the automaton $\mathcal{A}_{\overline{\mathbf{E}}}$ with $\overline{\mathbf{F}} = C_{f_j^k}(\mathbf{E})$ and $\overline{\mathbf{H}}_l = C_{g_i^l}(\mathbf{E})$ for $1 \leq l \leq m$; consequently $(\overline{\mathbf{H}}_1, \dots, \overline{\mathbf{H}}_m) \in g_i^{-1}(\overline{\mathbf{F}})$. From Proposition 6, it is equivalent to $(\mathbf{H}_1 \dots, \mathbf{H}_m) \in g^{-1}(\mathbf{F})$. Thus $(\mathbf{F}, g, \mathbf{H}_1 \dots, \mathbf{H}_m) = (h(C_{f_j^k}(\mathbf{E})), g, h(C_{g_i^l}(\mathbf{E})), \dots, h(C_{g_i^m}(\mathbf{E})))$ is a transition in the automaton $\mathcal{A}_{\mathbf{E}}$.

Since only equivalences are stated, $([C_{f_j{}^k}(\mathbf{E})], g, [C_{g_i^1}(\mathbf{E})] \dots, [C_{g_i^m}(\mathbf{E})])$ is a transition in $\mathfrak C$ if and only if $(\phi([C_{f_j{}^k}(\mathbf{E})]), g, \phi([C_{g_i^1}(\mathbf{E})]), \dots, \phi([C_{g_i^m}(\mathbf{E})]))$ is a transition in $\mathcal A_{\mathbf E}$.

Finally, for $c \in \Sigma_0$, $([C_{f_j}{}^k(\mathbf{E})], c)$ is a transition in $\mathfrak C$ if and only if $c \in \llbracket h(C_{f_j}{}^k(\mathbf{E})) \rrbracket$. Furthermore, it holds by construction that $(\phi([C_{f_j}{}^k(\mathbf{E})]), c)$ is a transition in $\mathcal A_{\mathbf E}$ if and only if $c \in \llbracket \phi([C_{f_j}{}^k(\mathbf{E})]) \rrbracket$. Consequently, $([C_{f_j}{}^k(\mathbf{E})], c)$ is a transition in $\mathcal C$ if and only if $(\phi([C_{f_j}{}^k(\mathbf{E})]), c)$ is a transition in $\mathcal A_{\mathbf E}$.

As a conclusion, ϕ is an isomorphism between \mathfrak{C} and \mathcal{A}_{E} .

4 Construction of the equation tree automaton $\mathcal{A}_{\rm E}$

In [8], the computation of the k-C-Continuations requires a preprocessing step which is the identification of subexpression of E in $O(|E|^2)$ time and space complexity. We propose an algorithm for the computation of the set of states with an O(|E|) time and space complexity.

4.1 Computation of the set of states $Q_{\overline{c}}/_{\sim_e}$

The main idea is to efficiently compute the quotient $\overline{C}_{\rm E}/_{\sim_e}$ by converting the syntax tree into a finite acyclic deterministic word automaton.

Let $T_{\rm E}$ be the syntax tree associated with E. The set of nodes of $T_{\rm E}$ is written as Nodes(E). For a node ν in Nodes(E), ${\rm sym}(\nu)$, father(ν), ${\rm son}(\nu)$, right(ν) and

left(ν) denote respectively the symbol, the father, the son, the right son and the left son of the node ν if they exist. We denote by E_{ν} the subexpression rooted at ν ; In this case we write ν_E to denote the node associated to E_{ν} . Let γ : Nodes(E) \cup { \bot } \rightarrow Nodes(E) \cup { \bot } be the function defined by:

$$\gamma(\nu) = \begin{cases} \text{father}(\nu) & \text{if sym(father}(\nu)) = *_c \text{ and } \nu \neq \nu_{\text{E}} \\ \text{right(father}(\nu)) & \text{if sym(father}(\nu)) = \cdot_c \\ \bot & \text{otherwise} \end{cases}$$

where \bot is an artificial node such that $\gamma(\bot) = \bot$. The ZPC-Structure is the syntax tree equipped with $\gamma(\nu)$ links. We extend the relation \preceq to the set of nodes of $T_{\rm E}$: For two nodes μ and ν we write $\nu \preceq \mu \Leftrightarrow T_{\rm E_{\nu}} \preceq T_{\rm E_{\mu}}$. We define the set $\Gamma_{\nu}(E) = \{\mu \in {\rm Nodes}(E) \mid \nu \preceq \mu \land \gamma(\mu) \neq \bot\}$ which is totally ordered by \preceq .

Proposition 7. Let E be linear, $1 \leq k \leq n$ be two integers and f be in $\Sigma_{\mathbb{E}} \cap \Sigma_n$. Then $C_{f^k}(E) = ((((E_{\nu_0} \cdot_{op(\nu_1)} E_{\gamma(\nu_1)}) \cdot_{op(\nu_2)} E_{\gamma(\nu_2)}) \cdot \cdots \cdot_{op(\nu_m)} E_{\gamma(\nu_m)})$ where ν_f is the node of $T_{\mathbb{E}}$ labelled by f, ν_0 is the k-child (ν_f) , $\Gamma_{\nu_f}(E) = \{\nu_1, \ldots, \nu_m\}$ and for $1 \leq i \leq m$, $op(\nu_i) = c$ such that $\operatorname{sym}(\operatorname{father}(\nu_i)) \in \{\cdot_c, *_c\}$.

Proof. By induction over the structure of E.

- 1. Let us suppose that $E = f(E_1, ..., E_n)$. Then $C_{f^k}(E) = E_k$. Since by definition ν_f is the root of T_E , k-child (ν_f) is the root of E_k . Hence $E_{\nu_0} = E_k = C_{f^k}(E)$.
- 2. Let us suppose that $E = g(E_1, \ldots, E_m)$ with $g \neq f$, or $E = E_1 + E_2$, or $E = E_1 \cdot_c E_2$ with $f \in \Sigma_{E_2}$. Then $C_{f^k}(E) = C_{f^k}(E_j)$ with $f \in \Sigma_{E_j}$. By induction hypothesis, $C_{f^k}(E_j) = ((((E_{\nu_0} \cdot_{op(\nu_1)} E_{\gamma(\nu_1)}) \cdot_{op(\nu_2)} E_{\gamma(\nu_2)}) \cdot \cdots \cdot_{op(\nu_m)} E_{\gamma(\nu_m)})$ where ν_f is the node of T_{E_j} labelled by f, ν_0 is the k-child (ν_f) , $\Gamma_{\nu_f}(E_j) = \{\nu_1, \ldots, \nu_m\}$ and for $1 \leq i \leq m$, $op(\nu_i) = c$ such that $\operatorname{sym}(\operatorname{father}(\nu_i)) \in \{\cdot_c, *_c\}$. Since $T_{E_j} \preccurlyeq T_E$, $C_{f^k}(E_j) = ((((E_{\nu_0} \cdot_{op(\nu_1)} E_{\gamma(\nu_1)}) \cdot_{op(\nu_2)} E_{\gamma(\nu_2)}) \cdot \cdots \cdot_{op(\nu_m)} E_{\gamma(\nu_m)})$ where ν_f is the node of T_E labelled by f, ν_0 is the k-child (ν_f) , $\Gamma_{\nu_f}(E_j) = \{\nu_1, \ldots, \nu_m\}$ and for $1 \leq i \leq m$, $op(\nu_i) = c$ such that $\operatorname{sym}(\operatorname{father}(\nu_i)) \in \{\cdot_c, *_c\}$.
- 3. Let us suppose that $E = E_1 \cdot_c E_2$ with $f \in \Sigma_{E_1}$ (resp. $E = E_1^{*c}$). Then $C_{f^k}(E) = C_{f^k}(E_1) \cdot_c G$ with $G \in \{E_1^{*c}, E_2\}$. By induction hypothesis, $C_{f^k}(E_1) = ((((E_{\nu_0} \cdot_{op(\nu_1)} E_{\gamma(\nu_1)}) \cdot_{op(\nu_2)} E_{\gamma(\nu_2)}) \cdot \cdots \cdot_{op(\nu_m)} E_{\gamma(\nu_m)})$ where ν_f is the node of T_{E_1} labelled by f, ν_0 is the k-child (ν_f) , $\Gamma_{\nu_f}(E_j) = \{\nu_1, \dots, \nu_m\}$ and for $1 \leq i \leq m$, $op(\nu_i) = c$ such that sym(father $(\nu_i)) \in \{\cdot_c, *_c\}$. Since $T_{E_1} \preccurlyeq T_E$, by setting $H = E_{\nu_{m+1}}$ and $op(\nu_{m+1})c$, $C_{f^k}(E_1) \cdot_c H = ((((E_{\nu_0} \cdot_{op(\nu_1)} E_{\gamma(\nu_1)}) \cdot_{op(\nu_2)} E_{\gamma(\nu_2)}) \cdot \cdots \cdot_{op(\nu_m)} E_{\gamma(\nu_m)}) \cdot_{op(\nu_{m+1})} E_{\gamma(\nu_{m+1})}$ where ν_f is the node of T_E labelled by f, ν_0 is the k-child (ν_f) , $\Gamma_{\nu_f}(E) = \{\nu_1, \dots, \nu_m, \nu_{m+1}\}$ and for $1 \leq i \leq m+1$, $op(\nu_i) = c$ such that sym(father $(\nu_i)) \in \{\cdot_c, *_c\}$.

Corollary 3. Let E be linear, $f \in (\Sigma_{E})_m$ and $k \leq m$. Then $|C_{f_i^k}(E)| \leq |E|^2$.

Example 2. Let Σ be the ranked alphabet such that $\Sigma_0 = \{a,b\}$, $\Sigma_1 = \{h\}$ and $\Sigma_2 = \{f\}$. Let $\mathbf{E} = (f(a,a) + f(a,a))^{*_a} \cdot_a h(b)$. Then $\overline{\mathbf{E}} = (f_1(a,a) + f_2(a,a))^{*_a} \cdot_a h_3(b)$. The ZPC-Structure associated with $\overline{\mathbf{E}}$ is represented in Figure 1 restricted to some γ links. As stated in Proposition 7, $C_{f_1^1}(\overline{\mathbf{E}}) = ((a \cdot_a (f_1(a,a) + f_2(a,a))^{*_a}) \cdot_a h_3(b)) = ((\mathbf{E}_{\nu_0} \cdot_a \mathbf{E}_{\gamma_{\nu_1}}) \cdot_a \mathbf{E}_{\gamma_{\nu_2}})$.

In order to identify the equivalent k-C-Continuations, we can sort them in lexicographic order. This can be done in $O(|E|^3)$ time and space complexity using Paige and Tarjan's Algorithm [12]. This is due to the fact that the size of k-C-Continuations is in $O(|E|^2)$ (by Corollary 3). This complexity has been improved by using k-Pseudo-Continuations instead of k-C-Continuations [3,7].

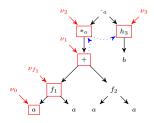


Fig. 1. ZPC-Structure of E.

A k-Pseudo-Continuation $l_{f_j^k}(E)$ of f_j in E is obtained from the k-C-Continuation $C_{f_j^k}(\overline{E})$ by replacing some subexpression \overline{F} of \overline{E} by a symbol $\psi(h(\overline{F}))$ such that for two subexpressions F and G of E: $\psi(F) = \psi(G) \Leftrightarrow F = G$.

Definition 5. Let H be a regular expression over Σ and ψ be a bijection that associates to each subexpression of E a symbol in an alphabet Ψ . We define the word $\psi'(H)$ over the alphabet $\Psi \cup \{\cdot_a \mid a \in \Sigma_0\}$ inductively as follows:

$$\psi'(\mathbf{H}) = \begin{cases} \psi'(\mathbf{F}) \cdot_{c} \psi(\mathbf{G}) & \text{if } \mathbf{H} = \mathbf{F} \cdot_{c} \mathbf{G} \text{ and } \mathbf{G} \text{ a subexpression of } \mathbf{E} \\ \psi(\mathbf{H}) & \text{if } \mathbf{H} \neq \mathbf{F} \cdot_{c} \mathbf{G} \text{ and } \mathbf{H} \text{ a subexpression of } \mathbf{E} \\ \varepsilon & \text{otherwise.} \end{cases}$$

The function ψ' is said to be an (E, Ψ) -encoding.

Definition 6. Let n and k be two integers such that $1 \leq k \leq n$, f_j be a symbol in $\operatorname{Pos}_{\mathbf{E}}(\mathbf{E})$ and ψ' an (\mathbf{E}, Ψ) -encoding for some alphabet Ψ . The k-Pseudo-Continuation of f_j in \mathbf{E} , denoted by $l_{f_j^k}(\mathbf{E})$, is the word over $\Psi \cup \{\cdot_a \mid a \in \Sigma_0\}$ defined by $l_{f_i^k}(\mathbf{E}) = \psi'(h(C_{f_i^k}(\overline{\mathbf{E}})))$.

In the following, we consider that the pseudo-continuations of E are defined over Ψ a finite subset of \mathbb{N} , bounded by the number of subexpressions of E.

Lemma 4. Let E and F be two regular expressions over an alphabet Σ such that E and F are two products of subexpressions of a regular expression H over Σ . Let ψ' be a (H, Ψ) -encoding. Then:

$$\psi'(E) = \psi'(F) \Leftrightarrow E = F.$$

Proof. Let us consider that ψ' is associated with the bijection ψ . Let us consider the possible roots of the expressions.

- 1. If the roots of E and F are notconcatenation products, $\psi'(E) = \psi'(F) \Leftrightarrow E = F$ since ψ is a bijection and $\psi'(E) = \psi(E)$ wedge $\psi'(F) = \psi(F)$.
- 2. Let us suppose that, without loss of generality, only the root of E is a concatenation product \cdot_c . Then the symbol \cdot_c appears in $\psi'(E)$ but not in $\psi'(F)$. Hence $\psi'(E) \neq \psi'(F)$ and $E \neq F$.
- 3. Finally, let us suppose that $E = E_1 \cdot_c E_2$ and $F = F_1 \cdot_d F_2$.
 - (a) If $E_2 \neq F_2$ then $\psi(E) \neq \psi(F)$ and then $\psi'(E)$ and $\psi'(F)$ do not end with the same symbol.

(b) Suppose that $E_2 = F_2$. If $c_c \neq d$, $\psi'(E)$ ends with $c_c \psi(E_2) \neq d \psi(E_2)$; Hence $\psi'(E) \neq \psi'(F)$ and $E \neq F$. Otherwise, by induction hypothesis, $\psi'(\mathbf{E}_1) = \psi'(F_1) \Leftrightarrow E_1 = F_1$. Hence $\psi'(\mathbf{E}_1) \cdot_c \psi(\mathbf{E}_2) = \psi'(F_1) \cdot_c \psi(F_2) \Leftrightarrow$ $E_1 \cdot_c E_2 = F_1 \cdot_d F_2.$

Proposition 8. The two following propositions hold:

- 1. $|l_{f_j^k}(\mathbf{E})|$ is at most linear w.r.t. $|\mathbf{E}|$, 2. $\sum_{f_j \in \text{Pos}_{\mathbf{E}}(\mathbf{E})_n, 1 \leq k \leq n} |\psi'(\mathbf{E}_{\nu_{f_j^k}})|$ is at most linear w.r.t. $|\mathbf{E}|$, with $\nu_{f_j^k} = k$ -child (ν_{f_j}) .

Proof. We define the function nbdot(E) as the number of left-associated concatenation operators in E as follows:

$$nbdot(E) = \begin{cases} nbdot(F) + 1 & \text{if } E = F \cdot_c G, \\ 0 & \text{otherwise.} \end{cases}$$

Let us first prove that $|\psi'(E)| \leq 2nbdot(E) + 1$. The proof proceeds by induction in the structure of E.

- 1. Whenever E is not a product, $\psi'(E) = \psi(E)$. Hence $|\psi'(E)| = 1$. Since nbdot(E) = 0, the condition is satisfied.
- 2. Suppose that $E = F \cdot_c G$. Hence

$$|\psi'(F \cdot_c G)| \le |\psi'(F)| + 2$$

 $\le 2(\text{nbdot}(F)) + 1 + 2$
 $= 2(\text{nbdot}(F) + 1) + 1$
 $= 2(\text{nbdot}(E)) + 1.$

Following Proposition 7, $|l_{f_i^k}(\mathbf{E})| \leq |\psi'(E_{\nu_0})| + 2m$ where ν_f is the node of $T_{\rm E}$ labelled by f, ν_0 is the k-child (ν_f) , $\Gamma_{\nu_f}({\rm E}) = \{\nu_1, \ldots, \nu_m\}$ and for $1 \leq 1$ $i \leq m, op(\nu_i) = c \text{ such that } \operatorname{sym}(\operatorname{father}(\nu_i)) \in \{\cdot_c, *_c\}. \text{ Since } \{\nu_1, \dots, \nu_m\}$ are ancestors of ν_0 , $|E_{\nu_0}| \leq |E| - m$. Consequently, $\operatorname{nbdot}(E_{\nu_0}) \leq |E| - m$. Moreover, from previous point, $|\psi'(E_{\nu_0})| \leq 2(\operatorname{nbdot}(E_{\nu_0})) + 1$. Consequently, $|l_{f_i^k}(\mathbf{E})| \le 2(|E| - m) + 1 + 2m = 2|E| + 1.$

Furthermore, since $|\psi'(\mathbf{E}_{\nu_{f^k}})| \leq 2(\mathrm{nbdot}(\mathbf{E}_{\nu_{f^k}})) + 1$, it holds:

$$\sum_{f_j \in \text{Pos}_{E}(E)_n} \sum_{1 \le k \le n} |\psi'(E_{\nu_{f_k^k}})| \le \sum_{f_j \in \text{Pos}_{E}(E)_n} \sum_{1 \le k \le n} 2(\text{nbdot}(E_{\nu_{f_k^k}})) + 1.$$

However, the concatenation operators below the node $\nu_{f_k^k}$ do not appears below another symbol. Consequently,

$$\sum_{f_j \in \operatorname{Pos_E}(\mathbf{E})_n} \sum_{1 \le k \le n} \operatorname{nbdot}(\mathbf{E}_{\nu_{f_i^k}}) \le |E|$$

Finally,

$$\sum_{f_j \in \operatorname{Pos_E}(\mathbf{E})_n} \sum_{1 \le k \le n} |\psi'(\mathbf{E}_{\nu_{f_j^k}})| \le 2|E| + |\Sigma_{\ge 1}|.$$

Proposition 9. Let $f_j \in \operatorname{Pos_E}(E)_n$, $g_i \in \operatorname{Pos_E}(E)_m$, $k \leq n$ and $p \leq m$ be two integers. Then $h(C_{f_i^k}(\overline{E})) = h(C_{g_i^p}(\overline{E})) \Leftrightarrow l_{f_i^k}(E) = l_{g_i^p}(E)$.

Proof. Direct Corollary of Lemma 4.

From Proposition 9 we can deduce that the k-C-Continuations identification can be achieved by considering the k-Pseudo-Continuations. In the following we show that this identification step (computation of \sim_e) can be done without the computation of the k-Pseudo-Continuations and that it amounts to the minimization of a word acyclic deterministic automaton. Before seeing how the identification of k-Pseudo-Continuations $l_{f_j^k}(E)$ is performed, we prove that the computation of the function ψ can be done in a linear time in the size of E.

Let us consider the syntax tree $T_{\rm E}$ associated with E. This syntax tree contains all the subexpressions of E. Each node ν in $T_{\rm E}$ corresponds to the subexpression E_{ν} of E. The equivalence relation \sim over the nodes of the tree $T_{\rm E}$ is defined by $\nu_1 \sim \nu_2 \Leftrightarrow E_{\nu_1} = E_{\nu_2}$. We show that the computation of the equivalence relation \sim amounts to the minimization of the word acyclic deterministic automaton $\mathcal{A}_{T_{\rm E}} = (Q, \Sigma_{\mathcal{A}}, \{\nu_{\rm E}\}, \{\nu_{T}\}, \delta)$, where ν_E is the node associated to the root of E, $Q = \operatorname{Nodes}(E) \cup \{\nu_T\} \cup \{\bot\}$ with $\nu_T, \bot \notin \operatorname{Nodes}(E)$, $\Sigma_{\mathcal{A}} = \Sigma_0 \cup \{g_+, d_+\} \cup \{*_a, g_{\cdot_a}, d_{\cdot_a} \mid a \in \Sigma_0\} \cup \{f^1, \dots, f^n \mid f \in \Sigma_n, n \geq 1\}$, and δ is defined by $\delta(\nu, *_a) = \operatorname{son}(\nu)$ if $\operatorname{sym}(\nu) = *_a, \delta(\nu, g_{\operatorname{sym}(\nu)}) = \operatorname{left}(\nu)$ and $\delta(\nu, d_{\operatorname{sym}(\nu)}) = \operatorname{right}(\nu)$ if $\operatorname{sym}(\nu) \in \{+, \cdot_a, a \in \Sigma_0\}$, $\delta(\nu, \operatorname{sym}(\nu)) = \nu_T$ if $\operatorname{sym}(\nu) \in \Sigma_0$, $\det(x, f^k) = k\operatorname{-child}(\nu)$ if $\operatorname{sym}(\nu) = f \in \Sigma_{\geq 1}$, and $\delta(\nu, x) = \bot$ in all otherwise.

Lemma 5. $E = F \Leftrightarrow \mathcal{L}(\mathcal{A}_{T_E}) = \mathcal{L}(\mathcal{A}_{T_E})$.

Proof. Let Σ_{A_E} (resp. Σ_{A_F}) be the alphabet of the automaton A_{T_E} (resp. A_{T_F}).

- 1. If E = F then $\mathcal{A}_{T_E} = \mathcal{A}_{T_F}$ and $\mathcal{L}(\mathcal{A}_{T_E}) = \mathcal{L}(\mathcal{A}_{T_F})$.
- 2. Suppose that $E \neq F$. Notice that any word w in $\mathcal{L}(\mathcal{A}_{T_E})$ (resp. $\mathcal{L}(\mathcal{A}_{T_F})$) starts with a symbol associated with the root of E (resp. F).
 - (a) Hence, if the roots of E and F are distinct, then $\mathcal{L}(\mathcal{A}_{T_E}) \cap \mathcal{L}(\mathcal{A}_{T_F}) = \emptyset$; Since $\mathcal{L}(\mathcal{A}_{T_E})$ is not empty by construction, $\mathcal{L}(\mathcal{A}_{T_E}) \neq \mathcal{L}(\mathcal{A}_{T_F})$.
 - (b) Otherwise, there exists an integer j such that $E = x(E_1, ..., E_k)$, $F = x(F_1, ..., F_k)$ and $E_j \neq F_j$. By induction over the size of E_j .
 - i. If $E_j \in \Sigma_0$, then since the roots are distincts, the word starting with the symbol associated to the node x followed by the symbol a is in $\mathcal{L}(\mathcal{A}_{T_E})$ but not in $\mathcal{L}(\mathcal{A}_{T_E})$.
 - ii. Otherwise, it holds by induction hypothesis that there exists a word in $\mathcal{L}(\mathcal{A}_{T_{E_j}})$ not in $\mathcal{L}(\mathcal{A}_{T_{F_j}})$. Hence there exists a word starting with a symbol associated to the node x followed by a word in $\mathcal{L}(\mathcal{A}_{T_{E_j}})$ that is in $\mathcal{L}(\mathcal{A}_{T_E})$ but not in $\mathcal{L}(\mathcal{A}_{T_F})$.

According to Lemma 5, $\nu_1 \sim \nu_2 \Leftrightarrow \mathcal{L}(\mathcal{A}_{T_{E_{\nu_1}}}) = \mathcal{L}(\mathcal{A}_{T_{E_{\nu_2}}})$, that is the equivalence relation \sim coincides with Myhill-Nerode equivalence [11] over the states of the automaton \mathcal{A}_{T_E} , that can be computed in O(|E|) time and space complexity using Revuz Algorithm [13].

Lemma 6. The computation of $\psi(F)$ for all subexpression F of E can be done in O(|E|) time and space complexity.

Proof. Let ν_1 and ν_2 be two nodes in T_E . As $\nu_1 \sim \nu_2 \Leftrightarrow E_{\nu_1} = E_{\nu_2} \Leftrightarrow \psi(E_{\nu_1}) =$ $\psi(E_{\nu_2})$, we can associate to each node ν in T_E (each E_{ν}) a symbol ($\psi(E_{\nu})$) which uniquely identifies its equivalence class $[\nu]_{\sim}$. Furthermore, according to Lemma 5, the computation of the equivalence relation \sim amounts to the minimization of the word acyclic deterministic automaton $\mathcal{A}_{T_{\mathbb{P}}}$, which can be performed in O(|E|)using Revuz Algorithm [13].

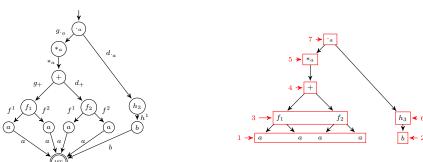


Fig. 3. The Equivalence Classes.

Fig. 2. The automaton $\mathcal{A}_{T_{\mathrm{E}}}$.

Example 3. Let us consider the regular expression $E = (f(a,a) + f(a,a))^{*a} \cdot a$ h(b) of the Example 2. Applying Myhill-Nerode equivalence [11] to the states of the automaton $\mathcal{A}_{T_{E}}$ (Figure 2) results in 7 equivalence classes labeled by $\Psi = \{1, 2, \dots, 7\}$. For example $\psi(f(a, a)) = 3$ and $\psi(E) = 7$ (Figure 3). Finally, $l_{f_1^1}(E) = 1 \cdot_a 6 \cdot_a 5.$

Recall that the k-Pseudo-Continuation identification can be achieved in $O(|E|^2)$ [4,8] using Paige and Tarjan's sorting algorithm [12]. In what follows we show that this step amounts to the minimization of the acyclic deterministic word automaton $\mathcal{B}_{T_{\overline{E}}} = (Q_{\mathcal{B}}, \Sigma_{\mathcal{B}}, \{\nu_T\}, \{\nu_{\overline{E}}\}, \delta_{\mathcal{B}})$ defined with $\nu_T \notin \text{Nodes}(\overline{E})$ and $\mathfrak{F} = \{f_j^k \mid 1 \leq k \leq m, f_j \in \operatorname{Pos_E}(E)_m\}$ by:

- $-Q_{\mathcal{B}} = (\operatorname{Nodes}(\overline{\mathbf{E}}) \setminus \Sigma_{0}) \cup \mathfrak{F} \cup \{\nu_{T}, \bot\},$ $-\Sigma_{\mathcal{B}} = \{\psi(\nu) \mid \nu \in \operatorname{Nodes}(\overline{\mathbf{E}}) \cap Q_{\mathcal{B}}\} \cup \mathfrak{F} \cup \{\cdot_{a} \mid a \in \Sigma_{0}\} \cup \{\varepsilon\},\$
- $-\delta_{\mathcal{B}}$ is defined as follows:

 - $\delta(\nu_T, f_j^i) = f_j^i$ for all $f_j^i \in \mathfrak{F}$, $\delta(f_j^k, \psi'(h(\mathbf{E}_{\nu_k}))) = f_j$ if ν_k is the k^{th} child of f_j , $\delta(\nu, \cdot_a \psi(\mathbf{E}_{\gamma(\nu)})) = \text{father}(\nu)$ if $\text{sym}(\text{father}(\nu)) \in \{\cdot_a, *_a\}$ and $\gamma(\nu) \neq \bot$, $\delta(\nu, \varepsilon) = \text{father}(\nu)$ and if $\gamma(\nu) = \bot$ and $\delta(\nu, x) = \bot$ in all otherwise.

Proposition 10. $\mathcal{L}(\mathcal{B}_{T_{\overline{E}}}) = \{ f_j^k \cdot l_{f_j^k}(E) \mid f_j \in \text{Pos}_E(E)_m, k \leq m \}$

Proof. By construction of $\mathcal{B}_{T_{\overline{E}}}$, there exists a path from any state f_i^k with $f_j \in \operatorname{Pos_E}(E)_m$ and $1 \leq k \leq m$ to the root of E labelled by $\psi'(E_{\nu_0}) \cdot_{op(\nu_1)}$ $\psi(\mathbf{E}_{\gamma(\nu_1)})\cdots \psi(\mathbf{E}_{\gamma(\nu_m)})$ where ν_{f_j} is the node of $T_{\overline{\mathbf{E}}}$ labelled by f_j , ν_0 is the k-child (ν_f) , $\Gamma_{\nu_{f_j}}(\mathbf{E}) = \{\nu_1, \dots, \nu_m\}$ and for $1 \leq i \leq m, op(\nu_i) = c$ such that $\operatorname{sym}(\operatorname{father}(\nu_i)) \in \{\cdot_c, *_c\}$. This word exactly corresponds to the word $\psi'(h(C_{f_i^k}(\overline{\mathbf{E}}))) = l_{f_i^k}(\mathbf{E}).$

Let f_j and g_i be two positions in Pos_E (E). As a direct consequence of Proposition 10, $C_{f_j^k}(\overline{\mathbb{E}}) \sim_e C_{g_i^p}(\overline{\mathbb{E}})$ if and only if the states f_j^k and g_i^p of $\mathcal{B}_{T_{\overline{\mathbb{E}}}}$ are equivalent. We eliminate the ε -transitions from the automaton $\mathcal{B}_{T_{\overline{\mathbb{E}}}}$. Since it has no ε -transitions cycles, this elimination can be performed in a linear time in the size of E. Hence, we obtain a more compacted but equivalent structure, which we denote by ε -free($\mathcal{B}_{T_{\overline{\mathbb{E}}}}$).

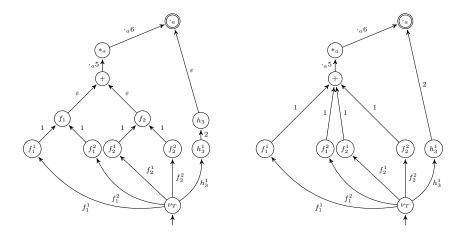


Fig. 4. The automaton $\mathcal{B}_{T_{\overline{\Sigma}}}$.

Fig. 5. The automaton ε -free $(\mathcal{B}_{T_{\overline{\Sigma}}})$.

The computation of the equivalence relation \sim_e can be performed by the computation of Myhill-Nerode relation [11] on the states of the automaton ε -free($\mathcal{B}_{T_{\overline{\mathbf{k}}}}$). This automaton is deterministic and acyclic.

Theorem 3. The relation \sim_e can be computed in O(|E|) time complexity.

Proof. The equivalence relation \sim_e coincides with Myhill-Nerode equivalence [11] on the states of the automaton ε -free $(\mathcal{B}_{T_{\overline{w}}})$.

This automaton is deterministic and acyclic and its size is linear with respect to |E| (Proposition 8). That can be computed in O(|E|) time and space complexity using Revuz Algorithm [13].

Example 4. Let us consider the regular expression $\mathbf{E} = (f(a,a) + f(a,a))^{*_a} \cdot_a h(b)$ of Example 2. The automaton $\mathcal{B}_{T_{\overline{\mathbf{E}}}}$ is represented by Figure 4. The automaton ε -free($\mathcal{B}_{T_{\overline{\mathbf{E}}}}$) is represented in Figure 5. Applying Myhill-Nerode equivalence to the automaton ε -free($\mathcal{B}_{T_{\overline{\mathbf{E}}}}$) results in the automaton in Figure 6. We deduce from this automaton that $C_{f_1^1}(\overline{\mathbf{E}}) \sim_e C_{f_2^2}(\overline{\mathbf{E}})$. Consequently the set of states of $\mathcal{C}_{\mathbf{E}}/\sim_e$ is $\{[C_{\varepsilon^1}(\overline{E})], [C_{f_1^1}(\overline{\mathbf{E}})], [C_{h_2^1}(\mathbf{E})]\}$.

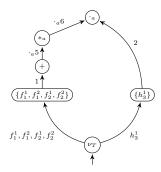


Fig. 6. The Minimal Automaton of ε -free($\mathcal{B}_{T_{\overline{n}}}$).

4.2 Computation of the set of transition rules

Using Proposition 3 and Proposition 4, we can show that the computation of the set of transitions of the equation tree automaton is performed by computing the function Follow. The computation of a transition rule using Proposition 3 requires a linear time, according to Proposition 2. Then for all transition rules we get an $O(|Q/\sim_e| \times |\mathbf{E}|)$ time and space complexity where Q is the set of k-C-Continuations of $\overline{\mathbf{E}}$. The computation of the set of states $Q_{\overline{C}}/\sim_e$ make possible the creation of non-coaccessible states. Removing these states requires an $O(|Q_{\overline{C}}/\sim_e| \cdot |\mathbf{E}|)$ time complexity.

Theorem 4. The equation tree automaton A_E of E can be computed in $O(|Q| \cdot |E|)$ time and space complexity with Q the set of states of A_E .

Proof. The equivalence relation \sim_e can be computed in $O(|\operatorname{E}|)$ time and space complexity and the set of transition rules can be performed by computing the function Follow. The computation of a transition rule using Proposition 3 requires a linear time, according to Proposition 2. Then for all transition rules we get an $O(|Q_{\overline{C}}/_{\sim_e}| \times |\operatorname{E}|)$ time and space complexity where $Q_{\overline{C}}$ is the set of k-C-Continuations of \overline{E} . Finally, removing not coaccessible states can be performed in linear time and results in the equation automaton.

5 A Full Example

Let $E = h(h(c,b)\cdot_c a, a)\cdot_b (f(a,h(c,b))\cdot_c a + g(a))^{*_b}$ be a regular expression defined over the ranked alphabet Σ such that $\Sigma^0 = \{a,b,c\}, \Sigma^1 = \{g\}, \Sigma^2 = \{f,h\}$ and $\overline{E} = h_1(h_2(c,b)\cdot_c a, a)\cdot_b (f_3(a,h_4(c,b))\cdot_c a + g_5(a))^{*_b}$ be its linearized form.

The computation of the k-C-Continuations of the E using the Definition 3 is given in Table 1.

$$\begin{array}{l} C_{h_{1}^{1}}(\overline{\mathbf{E}}) = \left(h_{2}(c,b) \cdot_{c} a\right) \cdot_{b} \left(f_{3}(a,h_{4}(c,b)) \cdot_{c} a + g_{5}(a)\right)^{*_{b}}, \\ C_{h_{1}^{2}}(\overline{\mathbf{E}}) = a \cdot_{b} \left(f_{3}(a,h_{4}(c,b)) \cdot_{c} a + g_{5}(a)\right)^{*_{b}}, \\ C_{f_{3}^{1}}(\overline{\mathbf{E}}) = a \cdot_{c} a \cdot_{b} \left(f_{3}(a,h_{4}(c,b)) \cdot_{c} a + g_{5}(a)\right)^{*_{b}}, \\ C_{f_{3}^{2}}(\overline{\mathbf{E}}) = \left(h_{4}(c,b) \cdot_{c} a\right) \cdot_{b} \left(f_{3}(a,h_{4}(c,b)) \cdot_{c} a + g_{5}(a)\right)^{*_{b}}, \\ C_{h_{2}^{1}}(\overline{\mathbf{E}}) = c \cdot_{c} a \cdot_{b} \left(f_{3}(a,h_{4}(c,b)) \cdot_{c} a + g_{5}(a)\right)^{*_{b}}, \\ C_{h_{2}^{2}}(\overline{\mathbf{E}}) = b \cdot_{c} a \cdot_{b} \left(f_{3}(a,h_{4}(c,b)) \cdot_{c} a + g_{5}(a)\right)^{*_{b}}, \\ C_{h_{4}^{1}}(\overline{\mathbf{E}}) = c \cdot_{c} a \cdot_{b} \left(f_{3}(a,h_{4}(c,b)) \cdot_{c} a + g_{5}(a)\right)^{*_{b}}, \\ C_{h_{4}^{2}}(\overline{\mathbf{E}}) = b \cdot_{c} a \cdot_{b} \left(f_{3}(a,h_{4}(c,b)) \cdot_{c} a + g_{5}(a)\right)^{*_{b}}, \\ C_{g_{5}^{1}}(\overline{\mathbf{E}}) = a \cdot_{b} \left(f_{3}(a,h_{4}(c,b)) \cdot_{c} a + g_{5}(a)\right)^{*_{b}}. \\ \mathbf{Table 1. The } k\text{-C-Continuations of E}. \end{array}$$

From Table 1, the Follow function can be computed (Table 2).

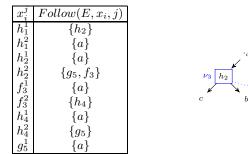


Table 2. The function Follow.

Fig. 7. The ZPC-Structure of E.

Finally, from Table 2, the transition function of C_E is the following:

any, from Table 2, the transition function of
$$C_{\mathbb{E}}$$
 is the following:
$$h(C_{h_{1}^{1}}(\overline{\mathbb{E}}), C_{h_{1}^{2}}(\overline{\mathbb{E}})) \to C_{\varepsilon^{1}}(\overline{\mathbb{E}}) \qquad h(C_{h_{1}^{1}}(\overline{\mathbb{E}}), C_{h_{2}^{2}}(\overline{\mathbb{E}})) \to C_{h_{1}^{1}}(\overline{\mathbb{E}})$$

$$= a \to C_{h_{1}^{2}}(\overline{\mathbb{E}}) \qquad f(C_{h_{1}^{1}}(\overline{\mathbb{E}}), C_{h_{2}^{2}}(\overline{\mathbb{E}})) \to C_{h_{2}^{1}}(\overline{\mathbb{E}})$$

$$= a \to C_{f_{3}^{1}}(\overline{\mathbb{E}}) \qquad h(C_{h_{4}^{1}}(\overline{\mathbb{E}}), C_{h_{4}^{2}}(\overline{\mathbb{E}})) \to C_{h_{2}^{2}}(\overline{\mathbb{E}})$$

$$= a \to C_{g_{5}^{1}}(\overline{\mathbb{E}}) \qquad h(C_{h_{4}^{1}}(\overline{\mathbb{E}}), C_{h_{4}^{2}}(\overline{\mathbb{E}})) \to C_{h_{4}^{2}}(\overline{\mathbb{E}})$$

$$= a \to C_{h_{4}^{1}}(\overline{\mathbb{E}})$$

$$= f(C_{f_{3}^{1}}(\overline{\mathbb{E}}), C_{f_{3}^{2}}(\overline{\mathbb{E}})) \to C_{h_{4}^{2}}(\overline{\mathbb{E}})$$

$$= g(C_{g_{5}^{1}}(\overline{\mathbb{E}})) \to C_{h_{4}^{2}}(\overline{\mathbb{E}})$$

The ZPC-structure associated to E is represented in Figure 7. The dotted links in Figure 7 represent the function γ :

$$\gamma(\nu_2) = \nu_5, \ \gamma(\nu_3) = \nu_4, \ \gamma(\nu_6) = \nu_5, \ \gamma(\nu_7) = \nu_8.$$

The automaton $\mathcal{A}_{T_{\mathrm{E}}}$ associated with E is represented in Figure 8.

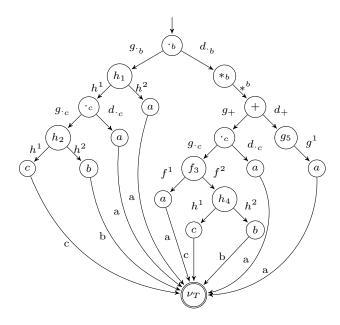


Fig. 8. The automaton $\mathcal{A}_{T_{\mathrm{E}}}$.

Applying Myhill-Nerode equivalence relation over the states of the automaton $\mathcal{A}_{T_{\mathrm{E}}}$ results in the automaton in Figure 9.

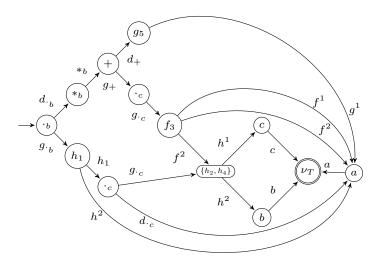


Fig. 9. The minimal automaton of $\mathcal{A}_{T_{\mathrm{E}}}$.

The computation of the equivalence relation \sim over the syntax tree associated to E is represented in the Figure 10. The number of equivalence classes in Figure 10 (12) corresponds exactly to the number of states of the minimal automaton of $\mathcal{A}_{T_{\rm E}}$. From these equivalence classes, we can define the ψ function (see Table 3).

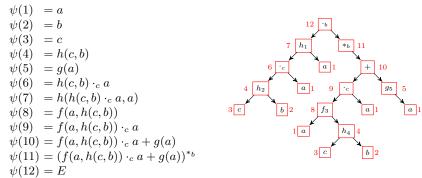


Table 3. The function ψ .

Fig. 10. The equivalence classes.

As we have seen, the computation of the equivalence relation \sim_e turns in the minimization of an acyclic deterministic automaton. The automaton $\mathcal{B}_{T_{\overline{E}}}$ associated with E is represented in Figure 11.

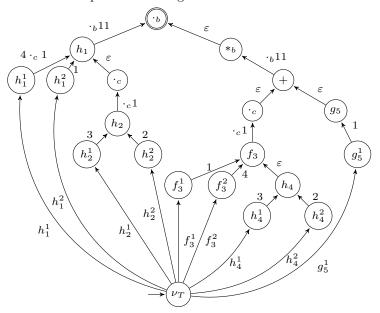


Fig. 11. The automaton $\mathcal{B}_{T_{\overline{w}}}$.

We eliminate the ε -transitions from the automaton $\mathcal{B}_{T_{\overline{E}}}$. Since this last has no ε -transitions cycles, this elimination can be performed in a linear time in the size of E. Hence, we obtain a structure which we denote ε -free($\mathcal{B}_{T_{\overline{E}}}$).

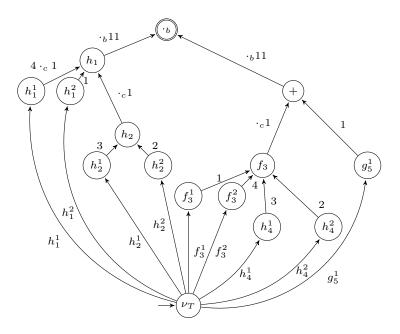


Fig. 12. The automaton ε -free $(\mathcal{B}_{T_{\overline{\Xi}}})$.

The computation of the equivalence relation \sim_e amounts to apply Myhill-Nerode relation on the states of the automaton ε -free $(\mathcal{B}_{T_{\overline{E}}})$. The result is represented in Figure 13.

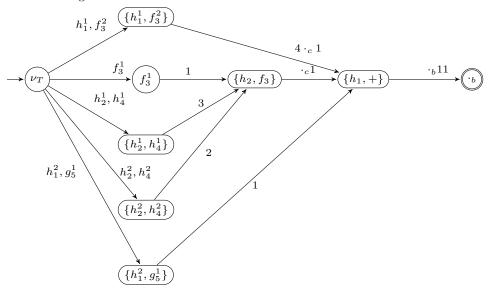


Fig. 13. The Minimal Automaton of ε -free $(\mathcal{B}_{T_{\overline{E}}})$.

The language recognized by $\mathcal{B}_{T_{\overline{\mathbb{E}}}}$ is the following:

$$\mathcal{L}(\mathcal{B}_{T_{\overline{E}}}) = \begin{cases} h_2^2, h_4^2 \} \cdot \{ 2 \cdot_c 1 \cdot_b 11 \} \\ \cup \{ h_2^1, h_4^1 \} \cdot \{ 3 \cdot_c 1 \cdot_b 11 \} \\ \cup \{ h_1^1, f_3^2 \} \cdot \{ 4 \cdot_c 1 \cdot_b 11 \} \\ \cup \{ h_1^2, g_5^1 \} \cdot \{ 1 \cdot_b 11 \} \\ \cup \{ f_3^1 1 \cdot_c 1 \cdot_b 11 \} \end{cases}$$

Let us notice that Proposition 9 is satisfied in Table 4.

\boldsymbol{x}	$xw \in \mathcal{L}(\mathcal{B}_{T_{\overline{\mathbf{E}}}})$	$C_x(\overline{E})$
h_1^1	$h_1^1 4 \cdot_c 1 \cdot_b 11$	$(h_2(c,b)\cdot_c a)\cdot_b (f_3(a,h_4(c,b))\cdot_c a+g_5(a))^{*_b}$
h_{1}^{2}	$h_1^2 1 \cdot_b 11$	$a \cdot_b (f_3(a, h_4(c, b)) \cdot_c a + g_5(a))^{*_b}$
h_2^1	$h_2^1 3 \cdot_c 1 \cdot_b 11$	$c \cdot_c a \cdot_b (f_3(a, h_4(c, b)) \cdot_c a + g_5(a))^{*_b}$
h_{2}^{2}	$h_2^2 2 \cdot_c 1 \cdot_b 11$	$b \cdot_c a \cdot_b (f_3(a, h_4(c, b)) \cdot_c a + g_5(a))^{*_b}$
f_{3}^{1}	$f_3^1 1 \cdot_c 1 \cdot_b 11$	$a \cdot_c a \cdot_b (f_3(a, h_4(c, b)) \cdot_c a + g_5(a))^{*_b}$
f_{3}^{2}	$f_3^2 4 \cdot_c 1 \cdot_b 11$	$(h_2(c,b)\cdot_c a)\cdot_b (f_3(a,h_4(c,b))\cdot_c a+g_5(a))^{*_b}$
h_4^1	$h_4^1 3 \cdot_c 1 \cdot_b 11$	$c \cdot_c a \cdot_b (f_3(a, h_4(c, b)) \cdot_c a + g_5(a))^{*_b}$
h_4^2	$h_4^2 2 \cdot_c 1 \cdot_b 11$	$b \cdot_c a \cdot_b (f_3(a, h_4(c, b)) \cdot_c a + g_5(a))^{*_b}$
g_5^1	$g_5^1 1 \cdot_b 11$	$a \cdot_b (f_3(a, h_4(c, b)) \cdot_c a + g_5(a))^{*_b}$

Table 4. $\mathcal{L}(\mathcal{B}_{T_{\overline{E}}})$ and k-C-Continuations.

Finally, the equation automaton \mathcal{A}_{E} associated with E is obtained from merging the states and the transitions using \sim_{e} . The transition function is:

$$\begin{split} h(\{C_{h_{1}^{1}}(\overline{\mathbf{E}}), C_{f_{3}^{2}}(\overline{\mathbf{E}})\}, \{C_{h_{1}^{2}}(\overline{\mathbf{E}}), C_{g_{5}^{1}}(\overline{\mathbf{E}})\}) &\to C_{\varepsilon^{1}}(\overline{\mathbf{E}}) \\ & a \to \{C_{h_{1}^{2}}(\overline{\mathbf{E}}), C_{g_{5}^{1}}(\overline{\mathbf{E}})\} \\ h(\{C_{h_{2}^{1}}(\overline{\mathbf{E}}), C_{h_{4}^{1}}(\overline{\mathbf{E}})\}, \{C_{h_{2}^{2}}(\overline{\mathbf{E}}), C_{h_{4}^{2}}(\overline{\mathbf{E}})\}) &\to \{C_{h_{1}^{1}}(\overline{\mathbf{E}}), C_{f_{3}^{2}}(\overline{\mathbf{E}})\} \\ & a \to \{C_{h_{2}^{1}}(\overline{\mathbf{E}}), C_{h_{4}^{1}}(\overline{\mathbf{E}})\} \\ & g(\{C_{g_{5}^{1}}(\overline{\mathbf{E}}), C_{h_{1}^{2}}(\overline{\mathbf{E}})\}) &\to \{C_{h_{2}^{2}}(\overline{\mathbf{E}}), C_{h_{4}^{2}}(\overline{\mathbf{E}})\} \\ & f(\{C_{f_{3}^{1}}(\overline{\mathbf{E}})\}, \{C_{f_{3}^{2}}(\overline{\mathbf{E}}), C_{h_{1}^{1}}(\overline{\mathbf{E}})\}) &\to \{C_{h_{2}^{2}}(\overline{\mathbf{E}}), C_{h_{4}^{2}}(\overline{\mathbf{E}})\} \\ & a \to \{C_{f_{3}^{1}}(\overline{\mathbf{E}})\} \end{split}$$

6 Conclusion

We presented a new and more efficient algorithm for the computation of the equation tree automaton from a regular tree expression by extending the notion of k-c-continuation from words to trees. We proved that a regular tree expression E can be converted into an equation tree automaton with an $O(|Q_{\overline{C}}/_{\sim_e}|\cdot|\operatorname{E}|)$ time and space complexity where Q is the set of k-C-Continuations of E.

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