

# On the Hardness of Almost–Sure Termination<sup>\*</sup>

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**Abstract.** This paper considers the computational hardness of computing expected outcomes and deciding (universal) (positive) almost–sure termination of probabilistic programs. It is shown that computing lower and upper bounds of expected outcomes is  $\Sigma_1^0$ – and  $\Sigma_2^0$ –complete, respectively. Deciding (universal) almost–sure termination as well as deciding whether the expected outcome of a program equals a given rational value is shown to be  $\Pi_2^0$ –complete. Finally, it is shown that deciding (universal) positive almost–sure termination is  $\Sigma_2^0$ –complete ( $\Pi_3^0$ –complete).

**Keywords:** probabilistic programs · expected outcomes · almost–sure termination · positive almost–sure termination · computational hardness

## 1 Introduction

Probabilistic programs [1] are imperative programs with the ability to toss a (possibly) biased coin and proceed their execution depending on the outcome of the coin toss. They are used in randomized algorithms, in security to describe cryptographic constructions (such as randomized encryption) and security experiments [2], and in machine learning to describe distribution functions that are analyzed using Bayesian inference [3]. Probabilistic programs are typically just a small number of lines, but hard to understand and analyze, let alone algorithmically. This paper considers the computational hardness of two main analysis problems (and variations thereof) for probabilistic programs:

1. Computing expected outcomes: Is the expected outcome of a program variable smaller than, equal to, or larger than a given rational number?
2. Deciding [universal] (positive) almost–sure termination: Does a program terminate [on all inputs] with probability one (within an expected finite number of computation steps)?

The first analysis problem is related to determining weakest pre–expectations of probabilistic programs [4, 5]. Almost–sure termination is an active field of research [6]. A lot of work has been done towards automated reasoning for almost–sure termination. For instance, [7] gives an overview of some particularly interesting examples of probabilistic logical programs and the according

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intuition for proving almost-sure termination. Arons *et al.* [8] reduce almost-sure termination to termination of non-deterministic programs by means of a planner. This idea has been further exploited and refined into a pattern-based approach with prototypical tool support [9].

Despite the existence of several (sometimes automated) approaches to tackle almost-sure termination, most authors claim that it must intuitively be harder than the termination problem for ordinary programs. To mention a few, Morgan [10] remarks that while partial correctness for small-scale examples is not harder to prove than for ordinary programs, the case for total correctness of a probabilistic loop must be harder to analyze. Esparza *et al.* [9] claim that almost-sure termination must be harder to decide than ordinary termination since for the latter a topological argument suffices while for the former arithmetical reasoning is needed. The computational hardness of almost-sure termination has however received scant attention. As a notable exception, [11] establishes that deciding almost-sure termination of certain concurrent probabilistic programs is in  $\Pi_2^0$ .

In this paper, we give precise classifications of the level of arithmetical reasoning that is needed to decide the aforementioned analysis problems by establishing the following results: We first show that computing lower bounds on the expected outcome of a program variable  $v$  after executing a probabilistic program  $P$  on a given input  $\eta$  is  $\Sigma_1^0$ -complete and therefore arbitrarily close approximations from below are computable. Computing upper bounds, on the other hand, is shown to be  $\Sigma_2^0$ -complete, thus arbitrarily close approximations from above are not computable in general. Deciding whether an expected outcome equals some rational is shown to be  $\Pi_2^0$ -complete.

For the second analysis problem—almost-sure termination—we obtain that deciding almost-sure termination of probabilistic program  $P$  on a given input  $\eta$  is  $\Pi_2^0$ -complete. While for ordinary programs we have a complexity leap when moving from the non-universal to the universal halting problem, we establish that *this is not the case for probabilistic programs*: Deciding universal a.s. termination turns out to be  $\Pi_2^0$ -complete too. The case for *positive* almost-sure termination is different however: While deciding (non-universal) positive almost-sure termination is  $\Sigma_2^0$ -complete, we show that universal positive almost-sure termination is  $\Pi_3^0$ -complete.

## 2 Preliminaries

As indicated, our hardness results will be stated in terms of levels in the arithmetical hierarchy—a concept we briefly recall:

**Definition 1 (Arithmetical Hierarchy [12, 13]).** *For every  $n \in \mathbb{N}$ , the class  $\Sigma_n^0$  is defined as  $\Sigma_n^0 = \{\mathcal{A} \mid \mathcal{A} = \{x \mid \exists y_1 \forall y_2 \exists y_3 \cdots \exists/\forall y_n : (x, y_1, y_2, y_3, \dots, y_n) \in \mathcal{R}\}$ ,  $\mathcal{R}$  is a decidable relation}, the class  $\Pi_n^0$  is defined as  $\Pi_n^0 = \{\mathcal{A} \mid \mathcal{A} = \{x \mid \forall y_1 \exists y_2 \forall y_3 \cdots \exists/\forall y_n : (x, y_1, y_2, y_3, \dots, y_n) \in \mathcal{R}\}$ ,  $\mathcal{R}$  is a decidable relation} and the class  $\Delta_n^0$  is defined as  $\Delta_n^0 = \Sigma_n^0 \cap \Pi_n^0$ . Note*

that we require that the values of the variables are drawn from a recursive domain. Multiple consecutive quantifiers of the same type can be contracted to one quantifier of that type, so the number  $n$  really refers to the number of necessary quantifier alternations rather than to the number of quantifiers used. A set  $\mathcal{A}$  is called **arithmetical**, iff  $\mathcal{A} \in \Gamma_n^0$ , for  $\Gamma \in \{\Sigma, \Pi, \Delta\}$  and  $n \in \mathbb{N}$ . The arithmetical sets form a strict hierarchy, i.e.  $\Delta_n^0 \subset \Gamma_n^0 \subset \Delta_{n+1}^0$  and  $\Sigma_n^0 \neq \Pi_n^0$  holds for  $\Gamma \in \{\Sigma, \Pi\}$  and  $n \geq 1$ . Furthermore, note that  $\Sigma_0^0 = \Pi_0^0 = \Delta_0^0 = \Delta_1^0$  is exactly the class of the decidable sets and  $\Sigma_1^0$  is exactly the class of the recursively enumerable sets.

Next, we recall the concept of many-one reducibility and completeness:

**Definition 2 (Many-One Reducibility and Completeness [13–15]).** Let  $\mathcal{A}, \mathcal{B}$  be arithmetical sets and let  $X$  be some appropriate universe such that  $\mathcal{A}, \mathcal{B} \subseteq X$ .  $\mathcal{A}$  is called **many-one-reducible** to  $\mathcal{B}$ , denoted  $\mathcal{A} \leq_m \mathcal{B}$ , iff there exists a computable function  $f: X \rightarrow X$ , such that  $\forall x \in X: (x \in \mathcal{A} \iff f(x) \in \mathcal{B})$ . If  $f$  is a function such that  $f$  many-one reduces  $\mathcal{A}$  to  $\mathcal{B}$ , we denote this by  $f: \mathcal{A} \leq_m \mathcal{B}$ . Note that  $\leq_m$  is transitive.

$\mathcal{A}$  is called  $\Gamma_n^0$ -**complete**, for  $\Gamma \in \{\Sigma, \Pi, \Delta\}$ , iff both  $\mathcal{A} \in \Gamma_n^0$  and  $\mathcal{A}$  is  $\Gamma_n^0$ -**hard**, meaning  $\mathcal{C} \leq_m \mathcal{A}$ , for any set  $\mathcal{C} \in \Gamma_n^0$ . Note that if  $\mathcal{A}$  is  $\Gamma_n^0$ -complete and  $\mathcal{A} \leq_m \mathcal{B}$ , then  $\mathcal{B}$  is necessarily  $\Gamma_n^0$ -hard. Furthermore, note that if  $\mathcal{A}$  is  $\Sigma_n^0$ -complete, then  $\mathcal{A} \in \Sigma_n^0 \setminus \Pi_n^0$ . Analogously if  $\mathcal{A}$  is  $\Pi_n^0$ -complete, then  $\mathcal{A} \in \Pi_n^0 \setminus \Sigma_n^0$ .

### 3 Probabilistic Programs

In order to speak about probabilistic programs and the computations performed by such programs, we briefly introduce the syntax and semantics we use:

**Definition 3 (Syntax).** Let  $\text{Var}$  be the set of program variables. The **set Prog of probabilistic programs** adheres to the following grammar:

$$\text{Prog} \longrightarrow v := e \mid \text{Prog}; \text{Prog} \mid \{\text{Prog}\} [p] \{\text{Prog}\} \mid \text{WHILE } (b) \{\text{Prog}\} ,$$

where  $v \in \text{Var}$ ,  $e$  is an arithmetical expression over  $\text{Var}$ ,  $p \in [0, 1] \subseteq \mathbb{Q}$ , and  $b$  is a Boolean expression over arithmetic expressions over  $\text{Var}$ . We call the set of programs that do not contain any probabilistic choices the **set of ordinary programs** and denote this set by **ordProg**.

The presented syntax is the one of the fully probabilistic<sup>1</sup> fragment of the probabilistic guarded command language (pGCL) originally due to McIver and Morgan [4]. We omitted **skip**-, **abort**-, and **if**-statements, as those are syntactic sugar. While assignment, concatenation, and the while-loop are standard programming constructs,  $\{P_1\} [p] \{P_2\}$  denotes a probabilistic choice between programs  $P_1$  (with probability  $p$ ) and  $P_2$  (with probability  $1 - p$ ). An operational semantics for pGCL programs is given below:

<sup>1</sup> Fully probabilistic programs may contain probabilistic but no non-deterministic choices.

**Definition 4 (Semantics).** Let the set of variable valuations be denoted by  $\mathbb{V} = \{\eta \mid \eta: \text{Var} \rightarrow \mathbb{Q}^+\}$ , let the set of program states be denoted by  $\mathbb{S} = (\text{Prog} \cup \{\downarrow\}) \times \mathbb{V} \times I \times \{L, R\}^*$ , for  $I = [0, 1] \cap \mathbb{Q}^+$ , let  $\llbracket e \rrbracket_\eta$  be the evaluation of the arithmetical expression  $e$  in the variable valuation  $\eta$ , and analogously let  $\llbracket b \rrbracket_\eta$  be the evaluation of the Boolean expression  $b$ . Then the **semantics of probabilistic programs** is given by the smallest relation  $\vdash \subseteq \mathbb{S} \times \mathbb{S}$  which satisfies the following inference rules:

$$\begin{array}{c}
\text{(assign)} \frac{}{\langle v := e, \eta, a, \theta \rangle \vdash \langle \downarrow, \eta[v \mapsto \max\{\llbracket e \rrbracket_\eta, 0\}], a, \theta \rangle} \\
\text{(concat1)} \frac{\langle P_1, \eta, a, \theta \rangle \vdash \langle P'_1, \eta', a', \theta' \rangle}{\langle P_1; P_2, \eta, a, \theta \rangle \vdash \langle P'_1; P_2, \eta', a', \theta' \rangle} \\
\text{(concat2)} \frac{}{\langle \downarrow; P_2, \eta, a, \theta \rangle \vdash \langle P_2, \eta, a, \theta \rangle} \\
\text{(prob1)} \frac{}{\langle \{P_1\} [p] \{P_2\}, \eta, a, \theta \rangle \vdash \langle P_1, \eta, a \cdot p, \theta \cdot L \rangle} \\
\text{(prob2)} \frac{}{\langle \{P_1\} [p] \{P_2\}, \eta, a, \theta \rangle \vdash \langle P_2, \eta, a \cdot (1 - p), \theta \cdot R \rangle} \\
\text{(while1)} \frac{\llbracket b \rrbracket_\eta = \text{True}}{\langle \text{WHILE } (b) \{P\}, \eta, a, \theta \rangle \vdash \langle P; \text{WHILE } (b) \{P\}, \eta, a, \theta \rangle} \\
\text{(while2)} \frac{\llbracket b \rrbracket_\eta = \text{False}}{\langle \text{WHILE } (b) \{P\}, \eta, a, \theta \rangle \vdash \langle \downarrow, \eta, a, \theta \rangle}
\end{array}$$

We use  $\sigma \vdash^k \tau$  in the usual sense.

The semantics is mostly straightforward except for two features: in addition to the program that is to be executed next and the current variable valuation, each state also stores a sequence  $\theta$  that encodes which probabilistic choices were made in the past (*Left* or *Right*) as well as the probability  $a$  that those choices were made. The graph that is spanned by the  $\vdash$ -relation can be seen as an unfolding of the Markov decision process semantics for pGCL provided by Gretz *et al.* [5] when restricting oneself to fully probabilistic programs.

## 4 Expected Outcomes and Termination Probabilities

In this section we formally define the notion of an expected outcome as well the notion of (universal) (positive) almost-sure termination. We start by investigating how state successors can be computed.

It is a well-known result due to Kleene that for any ordinary program  $P$  and a state  $\sigma$  the  $k$ -th successor of  $\sigma$  with respect to  $\vdash$  is unique and computable. If, however,  $P$  is a probabilistic program containing probabilistic choices, the  $k$ -th successor of a state need not be unique, because at various points of the execution

the program must choose a left or a right branch with some probability. However, if we resolve those choices by providing a sequence of symbols  $w$  over the alphabet  $\{L, R\}$  that encodes for all probabilistic choices which occur whether the *Left* or the *Right* branch shall be chosen at a branching point, we can construct a computable function that computes a unique  $k$ -th successor. Notice that for this purpose a sequence of finite length is sufficient. We obtain the following:

**Proposition 1 (The State Successor Function).** *Let  $\mathbb{S}_\perp = \mathbb{S} \cup \{\perp\}$ . There exists a total computable function  $T: \mathbb{N} \times \mathbb{S} \times \{L, R\}^* \rightarrow \mathbb{S}_\perp$ , such that for  $k \geq 1$*

$$T_0(\sigma, w) = \begin{cases} \sigma, & \text{if } w = \varepsilon, \\ \perp, & \text{otherwise,} \end{cases}$$

$$T_k(\sigma, w) = \begin{cases} T_{k-1}(\tau, w'), & \text{if } \sigma = \langle P, \eta, a, \theta \rangle \vdash \langle P', \eta', a', \theta \cdot b \rangle = \tau, \\ & \text{with } w = b \cdot w' \text{ and } b \in \{L, R, \varepsilon\}, \\ \perp & \text{otherwise.} \end{cases}$$

So  $T_k(\sigma, w)$  returns a successor state  $\tau$ , if  $\sigma \vdash^k \tau$ , whereupon exactly  $|w|$  inferences must use the (prob1)- or the (prob2)-rule and those probabilistic choices are resolved according to  $w$ . Otherwise  $T_k(\sigma, w)$  returns  $\perp$ . Note in particular that for both the inference of a terminal state  $\langle \downarrow, \eta, a, \theta \rangle$  within less than  $k$  steps as well as the inference of a terminal state through less or more than  $|w|$  probabilistic choices, the calculation of  $T_k(\sigma, w)$  will result in  $\perp$ . In addition to  $T$ , we will need two more computable operations for expressing expected outcomes, termination probabilities, and expected runtimes:

**Proposition 2.** *There exist two total computable functions  $\alpha: \mathbb{S}_\perp \rightarrow \mathbb{Q}^+$  and  $\wp: \mathbb{S}_\perp \times \text{Var} \rightarrow \mathbb{Q}^+$ , such that*

$$\alpha(\sigma) = \begin{cases} a, & \text{if } \sigma = \langle \downarrow, \_, a, \_ \rangle \\ 0, & \text{otherwise,} \end{cases} \quad \wp(\sigma, v) = \begin{cases} \eta(v) \cdot a, & \text{if } \sigma = \langle \downarrow, \eta, a, \_ \rangle \\ 0, & \text{otherwise,} \end{cases}$$

where  $\_$  represents an arbitrary value.

The function  $\alpha$  takes a state  $\sigma$  and returns the probability of reaching  $\sigma$ . The function  $\wp$  takes a state  $\sigma$  and a variable  $v$  and returns the probability of reaching  $\sigma$  multiplied with the value of  $v$  in the state  $\sigma$ . Both functions do that only if the provided state  $\sigma$  is a terminal state. Otherwise they return 0. Based on the above notions, we now define expected outcomes, termination probabilities and expected times until termination:

**Definition 5 (Expected Outcome, Termination Probability, and Expected Time until Termination).** *Let  $P \in \text{Prog}$ ,  $\eta \in \mathbb{V}$ ,  $v \in \text{Var}$ ,  $\sigma_{P,\eta} = \langle P, \eta, 1, \varepsilon \rangle$ , and for a finite alphabet  $A$  let  $A^{\leq k} = \bigcup_{i=0}^k A^i$ . Then*

1. the **expected outcome** of  $v$  after executing  $P$  on  $\eta$ , denoted  $\mathbf{E}_{P,\eta}(v)$ , is

$$\mathbf{E}_{P,\eta}(v) = \sum_{k=0}^{\infty} \sum_{w \in \{L, R\}^{\leq k}} \wp(T_k(\sigma_{P,\eta}, w), v),$$

2. the **probability that  $P$  terminates** on  $\eta$ , denoted  $\mathbf{Pr}_{P,\eta}(\downarrow)$ , is

$$\mathbf{Pr}_{P,\eta}(\downarrow) = \sum_{k=0}^{\infty} \sum_{w \in \{L, R\}^{\leq k}} \alpha(\mathbb{T}_k(\sigma_{P,\eta}, w)) ,$$

3. the **expected time until termination of  $P$  on  $\eta$** , denoted  $\mathbf{E}_{P,\eta}(\downarrow)$ , is

$$\mathbf{E}_{P,\eta}(\downarrow) = \sum_{k=0}^{\infty} \left( 1 - \sum_{w \in \{L, R\}^{\leq k}} \alpha(\mathbb{T}_k(\sigma_{P,\eta}, w)) \right) .$$

The expected outcome  $\mathbf{E}_{P,\eta}(v)$  as defined here coincides with the weakest pre-expectation  $wp.P.v(\eta)$  à la McIver and Morgan [4] for fully probabilistic programs. In the above definition for  $\mathbf{E}_{P,\eta}(v)$ , we sum over all possible numbers of inference steps  $k$  and sum over all possible sequences from length 0 up to length  $k$  for resolving all probabilistic choices. Using  $\wp$  we filter out the terminal states  $\sigma$  and sum up the values of  $\wp(\sigma, v)$ .

For the termination probability  $\mathbf{Pr}_P(\downarrow)$ , we basically do the same but we merely sum up the probabilities of reaching final states by using  $\alpha$  instead of  $\wp$ .

For the expected time until termination  $\mathbf{E}_{P,\eta}(\downarrow)$ , we go along the lines of [6]: It is stated there that the expected time until termination of  $P$  on  $\eta$  can be expressed as  $\sum_{k=0}^{\infty} \Pr(\text{“}P \text{ runs for more than } k \text{ steps on } \eta\text{”}) = \sum_{k=0}^{\infty} (1 - \Pr(\text{“}P \text{ terminates within } k \text{ steps on } \eta\text{”}))$ . We have expressed the latter in our set-up.

In order to investigate the complexity of calculating  $\mathbf{E}_{P,\eta}(v)$ , we define three sets:  $\mathcal{L}\mathcal{E}\mathcal{X}\mathcal{P}$ , which relates to the set of rational lower bounds of  $\mathbf{E}_{P,\eta}(v)$ ,  $\mathcal{R}\mathcal{E}\mathcal{X}\mathcal{P}$ , which relates to the set of rational upper bounds, and  $\mathcal{E}\mathcal{X}\mathcal{P}$  which relates to the value of  $\mathbf{E}_{P,\eta}(v)$  itself:

**Definition 6 ( $\mathcal{L}\mathcal{E}\mathcal{X}\mathcal{P}$ ,  $\mathcal{R}\mathcal{E}\mathcal{X}\mathcal{P}$ , and  $\mathcal{E}\mathcal{X}\mathcal{P}$ ).** *The sets  $\mathcal{L}\mathcal{E}\mathcal{X}\mathcal{P}$ ,  $\mathcal{R}\mathcal{E}\mathcal{X}\mathcal{P}$ ,  $\mathcal{E}\mathcal{X}\mathcal{P} \subset \text{Prog} \times \mathbb{V} \times \text{Var} \times \mathbb{Q}^+$  are defined as  $(P, \eta, v, q) \in \mathcal{L}\mathcal{E}\mathcal{X}\mathcal{P}$  iff  $q < \mathbf{E}_{P,\eta}(v)$ ,  $(P, \eta, v, q) \in \mathcal{R}\mathcal{E}\mathcal{X}\mathcal{P}$  iff  $q > \mathbf{E}_{P,\eta}(v)$ , and  $(P, \eta, v, q) \in \mathcal{E}\mathcal{X}\mathcal{P}$  iff  $q = \mathbf{E}_{P,\eta}(v)$ .*

Regarding the termination probability of a probabilistic program, the case of almost-sure termination is of special interest: We say that a program  $P$  *terminates almost-surely* on input  $\eta$  iff  $P$  terminates on  $\eta$  with probability 1. Furthermore, we say that  $P$  *terminates positively almost-surely* on  $\eta$  iff the expected time until termination of  $P$  on  $\eta$  is finite. Lastly, we say that  $P$  *terminates universally* (positively) *almost-surely*, if it does so on all possible inputs  $\eta$ . The problem of (universal) almost-sure termination can be seen as the probabilistic counterpart to the (universal) halting problem for ordinary programs.

In the following, we formally define the according problem sets:

**Definition 7 (Almost-Sure Termination Problem Sets).** *The sets  $\mathcal{A}\mathcal{S}\mathcal{T}$ ,  $\mathcal{P}\mathcal{A}\mathcal{S}\mathcal{T}$ ,  $\mathcal{U}\mathcal{A}\mathcal{S}\mathcal{T}$ , and  $\mathcal{P}\mathcal{A}\mathcal{S}\mathcal{T}$  are defined as follows:*

$$\begin{aligned} (P, \eta) \in \mathcal{A}\mathcal{S}\mathcal{T} &\iff \mathbf{Pr}_{P,\eta}(\downarrow) = 1 & (P, \eta) \in \mathcal{P}\mathcal{A}\mathcal{S}\mathcal{T} &\iff \mathbf{E}_{P,\eta}(\downarrow) < \infty \\ P \in \mathcal{U}\mathcal{A}\mathcal{S}\mathcal{T} &\iff \forall \eta: (P, \eta) \in \mathcal{A}\mathcal{S}\mathcal{T} & P \in \mathcal{U}\mathcal{P}\mathcal{A}\mathcal{S}\mathcal{T} &\iff \forall \eta: (P, \eta) \in \mathcal{P}\mathcal{A}\mathcal{S}\mathcal{T} \end{aligned}$$

*Notice that both  $\mathcal{P}\mathcal{A}\mathcal{S}\mathcal{T} \subset \mathcal{A}\mathcal{S}\mathcal{T}$  and  $\mathcal{U}\mathcal{P}\mathcal{A}\mathcal{S}\mathcal{T} \subset \mathcal{U}\mathcal{A}\mathcal{S}\mathcal{T}$  hold.*

## 5 The Hardness of Computing Expected Outcomes

In this section we investigate the computational hardness of deciding the sets  $\mathcal{L}\mathcal{E}\mathcal{X}\mathcal{P}$ ,  $\mathcal{R}\mathcal{E}\mathcal{X}\mathcal{P}$ , and  $\mathcal{E}\mathcal{X}\mathcal{P}$ . The first fact we establish is the  $\Sigma_1^0$ -completeness of  $\mathcal{L}\mathcal{E}\mathcal{X}\mathcal{P}$ . This result is established by reduction from the (non-universal) halting problem for ordinary programs:

**Theorem 1 (The Halting Problem [16]).** *The halting problem is a subset  $\mathcal{H} \subset \text{ordProg} \times \mathbb{V}$ , which is characterized as  $(P, \eta) \in \mathcal{H}$  iff  $\exists k \exists \eta' : \mathbb{T}_k(\sigma_{P,\eta}, \varepsilon) = \langle \downarrow, \eta', 1, \varepsilon \rangle$ . Let  $\overline{\mathcal{H}}$  denote the complement of the halting problem, i.e.  $\overline{\mathcal{H}} = (\text{ordProg} \times \mathbb{V}) \setminus \mathcal{H}$ .  $\mathcal{H}$  is  $\Sigma_1^0$ -complete and  $\overline{\mathcal{H}}$  is  $\Pi_1^0$ -complete.*

**Theorem 2.**  *$\mathcal{L}\mathcal{E}\mathcal{X}\mathcal{P}$  is  $\Sigma_1^0$ -complete.*

*Proof.* For showing  $\mathcal{L}\mathcal{E}\mathcal{X}\mathcal{P} \in \Sigma_1^0$ , observe that  $(P, \eta, v, q) \in \mathcal{L}\mathcal{E}\mathcal{X}\mathcal{P}$  iff  $\exists y : q < \sum_{k=0}^y \sum_{w \in \{L, R\}^{\leq k}} \wp(\mathbb{T}_k(\sigma_{P,\eta}, w), v)$ , which is a  $\Sigma_1^0$ -formula. Figure 1 (left) gives an intuition on this formula. For establishing  $\Sigma_1^0$ -hardness we use a reduction function  $f : \mathcal{H} \leq_m \mathcal{L}\mathcal{E}\mathcal{X}\mathcal{P}$  with  $f(Q, \eta) = (P, \eta, v, 1/2)$ , where  $P$  is the program  $v := 0; \{v := 1\}^{[1/2]}\{TQ; v := 1\}$  and  $TQ$  is an ordinary program that simulates  $Q$  on  $\eta$ . For details see [17].  $\square$

Theorem 2 implies that  $\mathcal{L}\mathcal{E}\mathcal{X}\mathcal{P}$  is recursively enumerable. This means that all lower bounds for expected outcomes can be effectively enumerated by some algorithm. Now, if upper bounds were recursively enumerable as well, then expected outcomes would be computable reals. However, the contrary will be shown by establishing that  $\mathcal{R}\mathcal{E}\mathcal{X}\mathcal{P}$  is  $\Sigma_2^0$ -complete, thus  $\mathcal{R}\mathcal{E}\mathcal{X}\mathcal{P} \notin \Sigma_1^0$  and hence  $\mathcal{R}\mathcal{E}\mathcal{X}\mathcal{P}$  is not recursively enumerable.  $\Sigma_2^0$ -hardness will be established by a reduction from the complement of the universal halting problem for ordinary programs:

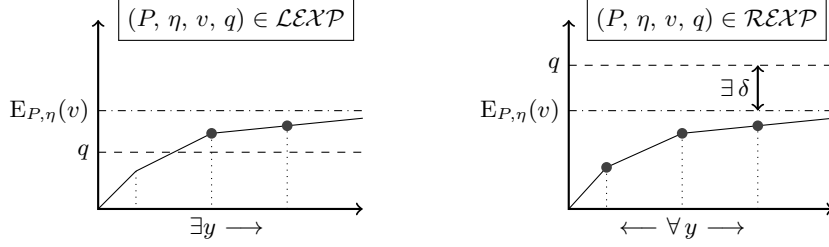
**Theorem 3 (The Universal Halting Problem [16]).** *The universal halting problem is a subset  $\mathcal{U}\mathcal{H} \subset \text{ordProg}$ , which is characterized as  $P \in \mathcal{U}\mathcal{H}$  iff  $\forall \eta : (P, \eta) \in \mathcal{H}$ . Let  $\overline{\mathcal{U}\mathcal{H}}$  denote the complement of  $\mathcal{U}\mathcal{H}$ , i.e.,  $\overline{\mathcal{U}\mathcal{H}} = \text{ordProg} \setminus \mathcal{U}\mathcal{H}$ .  $\mathcal{U}\mathcal{H}$  is  $\Pi_2^0$ -complete and  $\overline{\mathcal{U}\mathcal{H}}$  is  $\Sigma_2^0$ -complete.*

**Theorem 4.**  *$\mathcal{R}\mathcal{E}\mathcal{X}\mathcal{P}$  is  $\Sigma_2^0$ -complete.*

*Proof.* For showing  $\mathcal{R}\mathcal{E}\mathcal{X}\mathcal{P} \in \Sigma_2^0$ , observe that  $(P, \eta, v, q) \in \mathcal{R}\mathcal{E}\mathcal{X}\mathcal{P}$  iff  $\exists \delta \forall y : q - \delta > \sum_{k=0}^y \sum_{w \in \{L, R\}^{\leq k}} \wp(\mathbb{T}_k(\sigma_{P,\eta}, w), v)$ , which is a  $\Sigma_2^0$ -formula. Figure 1 (right) gives an intuition on this formula. For establishing  $\Sigma_2^0$ -hardness we use a reduction function  $f : \overline{\mathcal{U}\mathcal{H}} \leq_m \mathcal{R}\mathcal{E}\mathcal{X}\mathcal{P}$  with  $f(Q) = (P, \eta, v, 1)$ , where  $\eta$  is arbitrary but fixed and  $P$  is the probabilistic program

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i := 0; {c := 0} [0.5] {c := 1};
while (c ≠ 0) {i := i + 1; {c := 0} [0.5] {c := 1}};
k := 0; {c := 0} [0.5] {c := 1};
while (c ≠ 0) {k := k + 1; {c := 0} [0.5] {c := 1}};
v := 0; TQ ,
    
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**Fig. 1.** Schematic depiction of the formulae defining  $\mathcal{LEXP}$  and  $\mathcal{REXP}$ , respectively. In each diagram, the solid line represents the monotonically increasing graph of  $\sum_{k=0}^y \sum_{w \in \{L, R\}^{\leq k}} \wp(\mathbb{T}_k(\sigma_{P,\eta}, w), v)$  plotted over increasing  $y$ .

where  $TQ$  is a program that assigns the value  $2^{k+1}$  to the variable  $v$  if and only if  $Q$  halts on input  $g_Q(i)$  after exactly  $k$  steps (otherwise it assigns 0 to  $v$ ) and  $g_Q: \mathbb{N} \rightarrow \mathbb{V}$  is a computable bijection, such that  $\forall z \in \text{Var}: (g_Q(i))(z) \neq 0$  implies that  $z$  occurs in  $Q$ . For details see [17].  $\square$

Finally, we establish the following result regarding exact expected outcomes:

**Theorem 5.**  $\mathcal{EXP}$  is  $\Pi_2^0$ -complete.

*Proof.* For  $\mathcal{EXP}$  we can construct a  $\Pi_2^0$ -formula from the two formulae defining  $\mathcal{LEXP}$  and  $\mathcal{REXP}$ . For establishing  $\Pi_2^0$ -hardness we use the same reduction function  $f$  from the proof of Theorem 4 since for that function it holds that  $f: \mathcal{UH} \leq_m \mathcal{EXP}$ . For details see [17].  $\square$

## 6 The Hardness of Deciding Probabilistic Termination

This section presents the main contributions of this paper: Hardness results on several variations of almost-sure termination problems. We first establish that deciding almost-sure termination of a program on a given input is  $\Pi_2^0$ -complete:

**Theorem 6.**  $\mathcal{AST}$  is  $\Pi_2^0$ -complete.

*Proof.* For proving  $\mathcal{AST} \in \Pi_2^0$ , we show  $\mathcal{AST} \leq_m \mathcal{EXP}$  using the reduction function  $f: \mathcal{AST} \leq_m \mathcal{EXP}$  with  $f(Q, \eta) = (P, \eta, v, 1)$ , where  $v$  does not occur in  $Q$  and  $P$  is the program  $v := 0; Q; v := 1$ .

For establishing  $\Pi_2^0$ -hardness we use a reduction function  $f': \mathcal{UH} \leq_m \mathcal{AST}$  with  $f'(Q) = (P', \eta)$ , where  $\eta$  is arbitrary but fixed and  $P'$  is the program

```

i := 0; {c := 0} [0.5] {c := 1};
while (c ≠ 0) {i := i + 1; {c := 0} [0.5] {c := 1}};
SQ ,

```

where  $SQ$  is an ordinary program that simulates  $Q$  on  $g_Q(i)$  and  $g_Q: \mathbb{N} \rightarrow \mathbb{V}$  is the bijection from the proof of Theorem 4. For details see [17].  $\square$



While for ordinary programs there is a complexity leap when moving from the halting problem for some given input to the universal halting problem, we establish that *there is no such leap in the probabilistic setting*, i.e.  $\mathcal{UAST}$  is as hard as  $\mathcal{AST}$ :

**Theorem 7.**  $\mathcal{UAST}$  is  $\Pi_2^0$ -complete.

*Proof.* For showing  $\mathcal{UAST} \in \Pi_2^0$ , consider that by Theorem 6 there must exist a decidable relation  $\mathcal{R}$  such that  $(P, \eta) \in \mathcal{AST}$  iff  $\forall y_1 \exists y_2: (y_1, y_2, P, \eta) \in \mathcal{R}$ . By that we have that  $P \in \mathcal{UAST}$  iff  $\forall \eta \forall y_1 \exists y_2: (y_1, y_2, P, \eta) \in \mathcal{R}$ , which is a  $\Pi_2^0$ -formula.

It remains to show that  $\mathcal{UAST}$  is  $\Pi_2^0$ -hard. This can be done by proving  $\mathcal{AST} \leq_m \mathcal{UAST}$  as follows: On input  $(Q, \eta)$  the reduction function  $f: \mathcal{AST} \leq_m \mathcal{UAST}$  computes a probabilistic program  $P$  that first initializes all variables according to  $\eta$  and then executes  $Q$ .  $\square$

We now investigate the computational hardness of deciding *positive* almost-sure termination: It turns out that deciding  $\mathcal{PAST}$  is  $\Sigma_2^0$ -complete. Thus,  $\mathcal{PAST}$  becomes semi-decidable when given access to an  $\mathcal{H}$ -oracle whereas  $\mathcal{AST}$  does not. We establish  $\Sigma_2^0$ -hardness by a reduction from  $\overline{\mathcal{UH}}$ . This result is particularly counterintuitive as it means that for each ordinary program that *does not halt* on all inputs, we can *compute* a probabilistic program that *does halt* within an expected finite number of steps.

**Theorem 8.**  $\mathcal{PAST}$  is  $\Sigma_2^0$ -complete.

*Proof.* For showing  $\mathcal{PAST} \in \Sigma_2^0$ , observe that  $(P, \eta) \in \mathcal{PAST}$  iff  $\exists c \forall \ell: c > \sum_{k=0}^{\ell} \left(1 - \sum_{w \in \{L, R\}^{\leq k}} \alpha(\mathbb{T}_k(\sigma_{P, \eta}, w))\right)$ , which is a  $\Sigma_2^0$ -formula. For more details on this formula see [17].

It remains to show that  $\mathcal{PAST}$  is  $\Sigma_2^0$ -hard. For that we use a reduction function  $f: \overline{\mathcal{UH}} \leq_m \mathcal{PAST}$  with  $f(Q) = (P, \eta)$ , where  $\eta$  is arbitrary but fixed and  $P$  is the program

```

c := 1; i := 0; x := 0; term := 0; InitQ;
while (c ≠ 0){
  StepQ; if (term = 1){ Cheer; i := i + 1; term := 0; InitQ}
  {c := 0} [0.5] {c := 1}; x := x + 1 } ,
    
```

where  $InitQ \in \text{ordProg}$  is a program that initializes a simulation of the program  $Q$  on input  $g_Q(i)$  (recall the bijection  $g_Q: \mathbb{N} \rightarrow \mathbb{V}$  from Theorem 4),  $StepQ \in \text{ordProg}$  is a program that does one single (further) step of that simulation and sets **term** to 1 if that step has led to termination of  $Q$ , and  $Cheer \in \text{ordProg}$  is a program that executes  $2^x$  many effectless steps. In the following we refer to this as “cheering”<sup>2</sup>.

*Correctness of the reduction:* Intuitively, the program  $P$  starts by simulating  $Q$  on input  $g_Q(0)$ . During the simulation, it—figuratively speaking—gradually

<sup>2</sup> The program  $P$  cheers as it was able to prove the termination of  $Q$  on input  $g_Q(i)$ .

looses interest in further simulating  $Q$  by tossing a coin after each simulation step to decide whether to continue the simulation or not. If eventually  $P$  finds that  $Q$  has halted on input  $g_Q(0)$ , it “cheers” for a number of steps exponential in the number of coin tosses that were made so far, namely for  $2^x$  steps.  $P$  then continues with the same procedure for the next input  $g_Q(1)$ , and so on.

The variable  $x$  keeps track of the number of loop iterations (starting from 0), which equals the number of coin tosses. The  $x$ -th loop iteration takes place with probability  $1/2^x$ . One loop iteration consists of a constant number of steps  $c_1$  in case  $Q$  did not halt on input  $g_Q(i)$  in the current simulation step. Such an iteration therefore contributes  $c_1/2^x$  to the expected runtime of the probabilistic program  $P$ . In case  $Q$  did halt, a loop iteration takes a constant number of steps  $c_2$  plus  $2^x$  additional “cheering” steps. Such an iteration therefore contributes  $c_2+2^x/2^x = c_2/2^x + 1 > 1$  to the expected runtime. Overall, the expected runtime of the program  $P$  roughly resembles a geometric series with exponentially decreasing summands. However, for each time the program  $Q$  halts on an input, a summand of the form  $c_2/2^x + 1$  appears in this series. There are now two cases:

(1)  $Q \in \overline{UH}$ , so there exists some input  $\eta$  with minimal  $i$  such that  $g_Q(i) = \eta$  on which  $Q$  does not terminate. In that case, summands of the form  $c_2/2^x + 1$  appear only  $i - 1$  times in the series and therefore, the series converges—the expected time until termination is finite, so  $(P, \eta) \in \mathcal{PAST}$ .

(2)  $Q \notin \overline{UH}$ , so  $Q$  terminates on every input. In that case, summands of the form  $c_2/2^x + 1$  appear infinitely often in the series and therefore, the series diverges—the expected time until termination is infinite, so  $(P, \eta) \notin \mathcal{PAST}$ .  $\square$

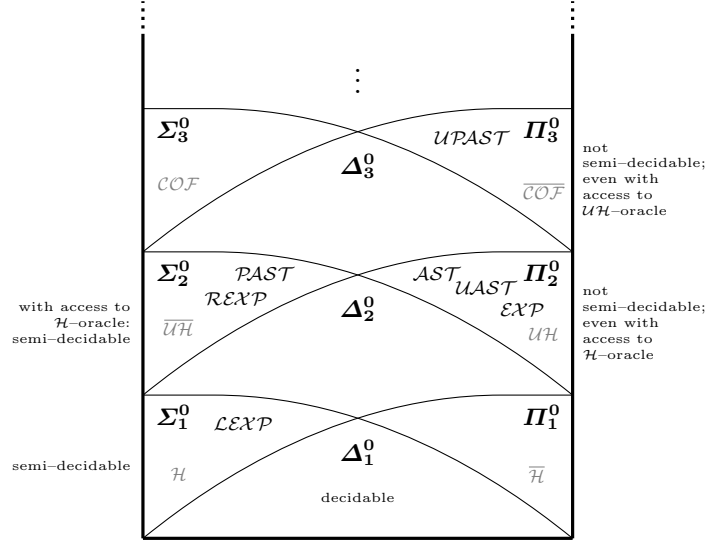
The final problem we study is *universal* positive almost-sure termination. In contrast to the non-positive version, we do have a complexity leap when moving from non-universal to universal positive almost-sure termination. We will establish that  $UPAST$  is  $\Pi_3^0$ -complete and thus even harder to decide than  $UAST$ . For the reduction, we make use of the following  $\Pi_3^0$ -complete problem:

**Theorem 9 (The Cofiniteness Problem [16]).** *The cofiniteness problem is a subset  $\mathcal{COF} \subset \text{ordProg}$ , which is characterized as  $P \in \mathcal{COF}$  iff  $\{\eta \mid (P, \eta) \in \mathcal{H}\}$  is cofinite. Let  $\overline{\mathcal{COF}}$  denote the **complement of  $\mathcal{COF}$** , i.e.  $\overline{\mathcal{COF}} = \text{ordProg} \setminus \mathcal{COF}$ .  $\mathcal{COF}$  is  $\Sigma_3^0$ -complete and  $\overline{\mathcal{COF}}$  is  $\Pi_3^0$ -complete.*

**Theorem 10.**  *$UPAST$  is  $\Pi_3^0$ -complete.*

*Proof.* By Theorem 8, there exists a decidable relation  $\mathcal{R}$  such that  $(P, \eta) \in \mathcal{PAST}$  iff  $\exists y_1 \forall y_2: (y_1, y_2, P, \eta) \in \mathcal{R}$ . Therefore  $UPAST$  is definable by  $P \in UPAST$  iff  $\forall \eta \exists y_1 \forall y_2: (y_1, y_2, P, \eta) \in \mathcal{R}$ , which gives a  $\Pi_3^0$ -formula.

It remains to show that  $UPAST$  is  $\Pi_3^0$ -hard. This is established by a reduction function  $f: \overline{\mathcal{COF}} \leq_m UPAST$  such that  $f(Q)$  gives nearly the same program as the reduction function from the proof of Theorem 8 except that in  $f(Q)$  the initialization  $i := 0$  is omitted. Thus, on input  $\eta$  the resulting program  $P$  also simulates  $Q$  successively on all inputs but starting from input  $g_Q(\eta(i))$  instead of  $g_Q(0)$ . For details, see [17].  $\square$



**Fig. 2.** The complexity landscape of determining expected outcomes and deciding (universal) (positive) almost-sure termination.

## 7 Conclusion

We have studied the computational complexity of solving a variety of natural problems which appear in the analysis of probabilistic programs: Computing lower bounds, upper bounds, and exact expected outcomes ( $LEXP$ ,  $REXP$ , and  $EXP$ ), deciding non-universal and universal almost-sure termination ( $AST$  and  $UAST$ ), and deciding non-universal and universal positive almost-sure termination ( $PAST$  and  $UPAST$ ). Our complexity results are summarized in Figure 2. All examined problems are complete for their respective level of the arithmetical hierarchy. We conjecture that all our results remain valid for programs with (minimizing) demonic non-determinism à la McIver & Morgan [4].

Future work consists of identifying program subclasses for which some of the studied problems become easier. One idea towards this would be to investigate the use of quantifier-elimination methods such as e.g. Skolemization.

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