Complexity of the Winner Determination Problem in Judgment Aggregation: Kemeny, Slater, Tideman, Young

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ABSTRACT

Judgment aggregation is a collective decision making framework where the opinions of a group of agents is combined into a collective opinion. This can be done using many different judgment aggregation procedures. We study the computational complexity of computing the group opinion for several of the most prominent judgment aggregation procedures. In particular, we show that the complexity of this winner determination problem for analogues of the Kemeny rule, the Slater rule and the Young rule lies at the \( \Theta^p_2 \)-level of the Polynomial Hierarchy (PH). Moreover, we show that the problem has a complexity at the \( \Sigma^p_3 \)-level of the PH for the analogue of Tideman’s procedure with a fixed tie-breaking rule, and at the \( \Sigma^p_2 \)-level of the PH for the analogue of Tideman’s procedure without a fixed tie-breaking rule.

Categories and Subject Descriptors

F.2 [Analysis of Algorithms and Problem Complexity]: General

General Terms

Theory

Keywords

Judgment Aggregation; Winner Determination; Complexity Theory; Bounded Query Complexity

1. INTRODUCTION

Collective decision making is central in the area of multiagent systems [6]. Judgment aggregation is a collective decision making framework that can be used for many applications [12, 17, 23]. In judgment aggregation, the goal is to combine the opinions of a group of individuals (or agents) on a set of propositions into a collective opinion reflecting the views of the group as a whole. As such, judgment aggregation generalizes the setting of preference aggregation, where the opinions are restricted to preferences over a given domain of alternatives [9]. There are many ways (or procedures) to combine the individual opinions, and choosing between such procedures involves deciding between various desirable properties that such procedures can have. Some of the most salient desiderata for these procedures are that they are consistent (that is, the resulting group opinion is a tenable position) and complete (that is, for each proposition, the resulting group opinion takes a clear position).

Another important property of judgment aggregation procedures concerns their computational complexity, i.e., the amount of time it takes to compute the group opinion. In this paper, expanding on previous work [13, 21], we study the computational complexity of the winner determination problem for several of the most prominent judgment aggregation procedures that are complete and consistent. In particular, we study the judgment aggregation analogues of the Kemeny rule, the Slater rule and the Young rule, as well as two variants of Tideman’s Ranked-Pairs rule (whose judgment aggregation analogue we call the Ranked-Agenda rule), all familiar from voting theory and preference aggregation [5].

Concretely, we study two different computational problems for the various procedures, one decision problem and one search problem. We consider the problem of deciding whether there exists an outcome that is deemed acceptable by the judgment aggregation procedure and that satisfies a number of additional requirements. In addition, we consider the computational task of producing one such acceptable outcome, if it exists. We argue that this latter formalization of the problem more adequately models the relevant computational properties of the various procedures. We show that the complexity (of both the decision and the search problem) for the Kemeny, Slater and Young rules lies at the \( \Theta^p_2 \)-level of the Polynomial Hierarchy (PH). Interestingly, these rules have exactly the same complexity in the setting of preference aggregation. In addition, we show that the complexity of Tideman’s Ranked-Agenda procedure lies at the \( \Delta^p_2 \)-level of the PH, in case of a fixed tie-breaking rule, and lies at the \( \Sigma^p_2 \)-level of the PH in case ties can be broken in arbitrary ways. It is interesting that both variants of the Ranked-Agenda rule are of higher computational complexity than the other three rules we consider in the setting of judgment aggregation, whereas in the case of preference aggregation both variants are of lower complexity than these rules. An overview of the complexity results that we obtain in this paper can be found in Table 1.

Understanding the complexity of the winner determination problem is a first step to making these important, but generally highly intractable procedures amenable to practical use in multiagent systems and related fields requiring the consistent aggregation of information coming from several
autonomous agents, or more generally, independent sources of information. Determining the exact location of these problems in the PH is very useful for determining what algorithmic approaches are best-suited to solve the problems in practice. For problems at the $\Theta^p_2$-level and at the $\Delta^p_2$-level, for instance, the method of iterative SAT solving could be used, which is generally more efficient than solving methods for problems at the $\Sigma^p_2$-level.

Related Work.

The computational complexity of the winner determination problem has been studied before, by Endriss et al. [13] for several procedures including the Kemeny rule, and by Lang and Slavkovik [21] for several procedures, including the Slater rule, the Kemeny rule and the variant of the Ranked- Agenda procedure without a fixed tie-breaking rule. However, those previous works consider different formalizations of the computational task of winner determination. First, they study only decision problems (and no search problems) [13, 21]. Second, the problems studied by Lang and Slavkovik involve checking whether all outcomes satisfy a certain property, rather than some outcome [21]. The computational complexity of various other computational tasks in judgment aggregation has also been studied, including problems related to manipulation, bribery and control [1, 3, 4, 8, 13].

Structure of the Paper.

We begin, in Section 2, with reviewing notions from logic, judgment aggregation and complexity theory. Then, in Section 3, we formally define the computational problems that we use to capture the task of winner determination, and we analyze their computational complexity for the various judgment aggregation procedures. Finally, in Section 4, we conclude, and suggest directions for further research.

2. PRELIMINARIES

In this section we review relevant material on logic, judgment aggregation and complexity theory.

2.1 Propositional Logic

A literal is a propositional variable $x$ or a negated variable $\neg x$. For literals $l \in \{x, \neg x\}$, we let $\text{Var}(l) \equiv x$ denote the variable occurring in $l$. A clause is a finite set of literals, not containing a complementary pair $x, \neg x$, and is interpreted as the disjunction of these literals. We let $\bot$ denote the empty clause. A formula in conjunctive normal form (CNF) is a finite set of clauses, interpreted as the conjunction of these clauses. We define the size $||\varphi||$ of a CNF formula $\varphi$ to be $\sum_{c \in \varphi} |c|$: the number of clauses of $\varphi$ is denoted by $|\varphi|$. For a CNF formula $\varphi$, the set $\text{Var}(\varphi)$ denotes the set of all variables $x$ such that some clause of $\varphi$ contains $x$ or $\neg x$. We use the standard notion of (truth) assignments $\alpha : \text{Var}(\varphi) \rightarrow \{0, 1\}$ for Boolean formulas and truth of a formula under such an assignment. We let $\text{SAT}$ denote the problem of deciding whether a given propositional formula is satisfiable. For every propositional formula $\varphi$, we let $\neg \varphi$ denote the complement of $\varphi$, i.e., $\neg \varphi = \neg \psi$ if $\varphi$ is not of the form $\neg \psi$, and $\neg \varphi = \psi$ if $\varphi$ is of the form $\neg \psi$.

2.2 Judgment Aggregation

An agenda is a finite nonempty set $\Phi$ of propositional formulas that does not contain any doubly-negated formulas and that is closed under complementation. Moreover, if $\Phi = \{\varphi_1, \ldots, \varphi_m, \neg \varphi_1, \ldots, \neg \varphi_m\}$ is an agenda, then we let $[\Phi] = \{\varphi_1, \ldots, \varphi_m\}$ denote the preagenda associated with the agenda $\Phi$. A judgment set $J$ for an agenda $\Phi$ is a subset $J \subseteq \Phi$. We call a judgment set $J$ complete if either $\varphi \in J$ or $\neg \varphi \in J$, for all $\varphi \in \Phi$; we call it complement-free if for all $\varphi \in \Phi$ it is not the case that both $\varphi$ and $\neg \varphi$ are in $J$; and we call it consistent if there exists a truth assignment that makes all formulas in $J$ true.

In addition, we associate with each agenda $\Phi$ an integrity constraint $\Gamma$, that can be used to explicitly represent logical dependencies between agenda issues. Integrity constraints for agendas have been considered in previous literature [10, 21], and the notion bears resemblance to the framework of binary aggregation with integrity constraints [15, 16]. Such an integrity constraint $\Gamma$ consists of a single propositional formula. In the remainder of the paper, if no integrity constraint is specified, we implicitly assume that $\Gamma = \top$. We say that a judgment set $J$ is $\Gamma$-consistent if there exists a truth assignment that simultaneously makes all formulas in $J$ and $\Gamma$ true. Let $J(\Phi, \Gamma)$ denote the set of all complete and $\Gamma$-consistent subsets of $\Phi$.

Let $\mathcal{N}$ be a finite set of individuals (or agents). A judgment aggregation procedure (or rule) for the agenda $\Phi$ and the set $\mathcal{N}$ of individuals is a function $F$ that takes as input a profile $J \in J(\Phi, \Gamma)^n$, consisting of a sequence $(J_1, \ldots, J_n)$ of $n = |\mathcal{N}|$ complete and $\Gamma$-consistent judgment sets, and that produces a non-empty set of non-empty judgment sets, i.e., it produces an element in $F(\mathcal{P}(\mathcal{P}(\Phi)) \setminus \{\emptyset\}) \setminus \{\emptyset\}$. We call a judgment aggregation procedure $F$ resolute if for any profile $J$ it returns a singleton, i.e., $|F(J)| = 1$; otherwise, we call $F$ irresolute. An example of a resolute judgment aggregation procedure is the strict majority rule $F^\text{maj}$, where $\varphi \in F^\text{maj}(J)$ if and only if $\varphi$ occurs in the strict majority of judgment sets in $J$, for all $\varphi \in \Phi$. We call $F$ complete, complement-free and ($\Gamma$-)consistent, if $J$ is complete, complement-free and ($\Gamma$-)consistent, respectively, for every $J \in J(\Phi, \Gamma)^n$ and every $J \in F(J)$.

As an example, consider the following profile for a judgment aggregation problem with preagenda $\Phi = \{p, q, p \lor q\}$, integrity constraint $\Gamma = \neg (p \land q)$, and three individuals:

<table>
<thead>
<tr>
<th></th>
<th>$p$</th>
<th>$q$</th>
<th>$p \lor q$</th>
</tr>
</thead>
<tbody>
<tr>
<td>individual 1</td>
<td>no</td>
<td>yes</td>
<td>yes</td>
</tr>
<tr>
<td>individual 2</td>
<td>yes</td>
<td>no</td>
<td>yes</td>
</tr>
<tr>
<td>individual 3</td>
<td>no</td>
<td>no</td>
<td>no</td>
</tr>
<tr>
<td>majority</td>
<td>no</td>
<td>no</td>
<td>yes</td>
</tr>
</tbody>
</table>

Thus, respecting majorities leads to an inconsistent outcome for this profile. This is an instance of the well-known discursive dilemma [22]. In this paper, we will consider several judgment aggregation procedures that resolve this dilemma by always picking from the set of all consistent judgment sets. There currently is no consistent naming convention for these aggregation procedures in the literature. Here, to identify the procedures that we consider, we use names from well-known voting rules that most closely resemble them [5]. The procedures that we consider have been studied as judgment aggregation procedures before. We consider procedures that resemble the Kemeny rule [24, 25], the Slater rule [24], the Young rule [20], and the Ranked-Agenda rule [20, 26].

The former three aggregation procedures are based on a notion of score. For $F \in \{\text{Kemeny}, \text{Slater}, \text{Young}\}$, we define $\text{Score}_F$ as follows. Here $J$ denotes a single consistent and complete judgment set, $J \in J(\Phi, \Gamma)^n$ denotes
Table 1: Overview of complexity results.

<table>
<thead>
<tr>
<th>F</th>
<th>WinDet(F)</th>
<th>FWinDet(F)</th>
<th>location in the PH</th>
</tr>
</thead>
<tbody>
<tr>
<td>Kemeny</td>
<td>$P_{NP}^{log\text{-}complete}$</td>
<td>$FP_{NP}^{log\text{-}wit\text{-}complete}$</td>
<td>$\Theta_2^p$-level</td>
</tr>
<tr>
<td>Slater</td>
<td>$P_{NP}^{log\text{-}complete}$</td>
<td>$FP_{NP}^{log\text{-}wit\text{-}complete}$</td>
<td>$\Theta_2^p$-level</td>
</tr>
<tr>
<td>Young</td>
<td>$P_{NP}^{log\text{-}complete}$</td>
<td>$FP_{NP}^{log\text{-}wit\text{-}complete}$</td>
<td>$\Theta_2^p$-level</td>
</tr>
<tr>
<td>Tideman (fixed tie-breaking)</td>
<td>$P_{NP}^{complete}$</td>
<td>$FP_{NP}^{complete}$</td>
<td>$\Delta_2^p$-level</td>
</tr>
<tr>
<td>Tideman</td>
<td>$\Sigma_2^p$-complete</td>
<td>$\Sigma_2^p$-complete</td>
<td>$\Sigma_2^p$-level</td>
</tr>
</tbody>
</table>

2.3 Complexity Theory

Readers familiar with search problems and bounded query complexity may skip this section.

Search Problems.

In this paper, we will assume knowledge of to the well-known complexity classes P and NP, consisting of decision problems. In addition, we will consider search problems. Let $\Sigma$ be an alphabet. A search problem is a binary relation $R$ over strings in $\Sigma^*$. For any input string $x \in \Sigma^*$, we let $R(x) = \{ y \in \Sigma^* : (x,y) \in R \}$ denote the set of solutions for $x$. We say that a Turing machine $T$ solves $R$ if on input $x \in \Sigma^*$ the following holds: if there exists at least one $y$ such that $(x,y) \in R$, then $T$ accepts $x$ and outputs some $y$ such that $(x,y) \in R$; otherwise, $T$ rejects $x$.

With any search problem $R$ we associate a decision problem $S_R$, defined by $S_R = \{ x \in \Sigma^* : \text{there exists some } y \in \Sigma^* \text{ such that } (x,y) \in R \}$. We will use the following notion of reductions for search problems. A polynomial-time Levin reduction from one search problem $R_1$ to another search problem $R_2$ is a pair of polynomial-time computable functions $(g_1,g_2)$ such that (1) the function $g_1$ is a many-one reduction from $R_1$ to $S_{R_2}$, i.e., every $x \in \Sigma^*$ holds that $x \in S_{R_2}$ if and only if $g_1(x) \in S_{R_1}$; and (2) for every string $x \in S_{R_1}$ and every solution $y \in S_{R_2}(g_1(x))$ it holds that $(x,g_2(x,y)) \in R_1$. For more details, we refer to textbooks on the topic [14].

Complexity Classes.

The complexity class FP consists of those search problems that can be computed by a polynomial-time deterministic Turing machine, and the class FNP consists of those search problems that can be computed by a polynomial-time non-deterministic Turing machine.

Moreover, we will use complexity classes that are based on Turing machines that have access to an oracle. Let $C$ be a complexity class with decision problems. A Turing machine $T$ with access to a yes-no C oracle is a Turing machine with a dedicated oracle tape and dedicated states $q_{oracle}, q_{yes}$ and $q_{no}$. Whenever $T$ is in the state $q_{oracle}$, it does not proceed according to the transition relation, but instead it transitions into the state $q_{yes}$ if the oracle tape contains a string $x$ that is a yes-instance, and into the state $q_{no}$ if $x \not\in C$, and it transitions into the state $q_{no}$ if $x \not\in C$. Let $C$ be a complexity class with search problems. Similarly, a Turing machine with access to a witness $C$ oracle has a dedicated oracle tape and dedicated states $q_{oracle}, q_{yes}$ and $q_{no}$. Also, whenever $T$ is in the state $q_{oracle}$ it transitions into the state $q_{yes}$ if the oracle tape contains a string $x$ that there exists some $y$ such that $C(x,y)$, and in addition the contents of the oracle tape are replaced by (the encoding of) such an $y$; it transitions into the state $q_{no}$ if there exists no $y$ such that $C(x,y)$. Such transitions are called oracle queries.
We point out that the notion of algorithms that have access to witness FNP oracles accurately models algorithms that can call a SAT solver (modulo the running time of the SAT solver), as SAT solvers also return a satisfying assignment if it exists.

In this paper, we will consider the following complexity classes that are based on oracle machines. The class \( P^{\text{NP}}[\log] \) consists of all decision problems that can be decided by a deterministic polynomial-time Turing machine that has access to a yes-no NP oracle, and on any input of length \( n \) queries the oracle at most \( O(\log n) \) many times. This class coincides with the class \( P^{\text{NP}} \) and is also known as \( \Theta^p_2 \).

The class \( FP^{\text{NP}}[\log,\text{wit}] \) consists of all search problems that can be solved by a deterministic polynomial-time Turing machine that has access to a witness FNP oracle, and on any input of length \( n \) queries the oracle at most \( O(\log n) \) many times. In a sense, it is the search variant of \( P^{\text{NP}}[\log] \).

The class \( \Sigma^p_2 \) consists of all search problems that can be computed by a nondeterministic polynomial-time Turing machine that has access to a witness FNP oracle. In a sense, it is the search variant of \( P^{\text{NP}}[\log] \).

We say that a decision problem \( P \) is complete for a complexity class \( C \) containing decision problems if any problem \( Q \in C \) is polynomial-time (many-one) reducible to \( P \). Similarly, we say that a search problem \( R \) is complete for a complexity class \( C \) containing search problems if for any problem \( R' \in C \) there exists a polynomial-time Levin reduction from \( R' \) to \( R \).

**Complete Problems.**

Next, we consider a number of search and decision problems that are complete for the various complexity classes that we consider in this paper. The decision problem SAT is NP-complete. This problem consists of deciding whether a given propositional formula \( \varphi \) has a satisfying assignment. The corresponding search problem, FSAT, consisting of all pairs \((\varphi, \alpha)\), where \( \alpha \) is a satisfying assignment for \( \varphi \), is FNP-complete. Next, consider the following problems.

- **X-MAX-MODEL**
  
  **Input:** a formula \( \varphi \) and a subset \( X \subseteq \text{Var}(\varphi) \) of variables of \( \varphi \).

  **Output:** a model \( M \) of \( \varphi \) that sets a maximum number of variables in \( X \) to true, if such a model exists.

- **X-MAX-MODEL-PARITY**
  
  **Instance:** a formula \( \varphi \), a subset \( X \subseteq \text{Var}(\varphi) \) of variables of \( \varphi \), and a variable \( x_0 \in X \).

  **Question:** is there a model \( M \) of \( \varphi \) such that \( X \)-MAX-MODEL\((\varphi, M)\) and \( M \) sets \( x_0 \) to true?

The search problem \( X\text{-MAX-MODEL} \) is complete for \( FP^{\text{NP}}[\log,\text{wit}] \) [7]. Moreover, the decision problem \( X\text{-MAX-MODEL-PARITY} \) is complete for \( P^{\text{NP}}[\log] \). This follows from the fact that all predicates in \( P^{\text{NP}}[\log,\text{wit}] \) are computable in \( P^{\text{NP}}[\log] \) [18, Corollary 6.3.5].

- **FLEX-MAX-MODEL**
  
  **Input:** a formula \( \varphi \) with \( \text{Var}(\varphi) = \{x_1, \ldots, x_n\} \).

  **Output:** a lexicographically maximal model \( M \) of \( \varphi \), if such a model exists. (Here, the variables are ordered \( x_1 < \cdots < x_n \).

- **LEX-MAX-MODEL**
  
  **Instance:** a formula \( \varphi \) and a variable \( x_0 \in \text{Var}(\varphi) \).

  **Question:** is there a model \( M \) of \( \varphi \) such that \( \text{FLEX-MAX-MODEL}(\varphi, M) \) and \( M \) sets \( x_0 \) to true?

The search problem \( \text{FLEx-MAX-MODEL} \) is complete for \( FP^{\text{NP}} \), and the decision problem \( \text{LEX-MAX-MODEL} \) is complete for \( P^{\text{NP}} \) [19].

- **FQSAT**
  
  **Input:** a quantified Boolean formula \( \varphi = \exists X. \forall Y. \psi \), where \( \psi \) is quantifier-free.

  **Output:** an assignment \( \alpha \) to the variables in \( X \), such that \( \forall Y. \psi[\alpha] \) is true, if it exists.

The decision problem \( \text{QSAT} \) associated to \( \text{FQSAT} \) is well-known to be \( \Sigma^p_2 \)-complete. The search problem \( \text{FQSAT} \) is complete for \( \Sigma^p_2 \). This can be shown straightforwardly by modifying the proof of \( \Sigma^p_2 \)-completeness for \( \text{QSAT} \) [27, 28], which can also be found in many textbooks (e.g., [2]).

3. **COMPLEXITY RESULTS**

For the various judgment aggregation procedures \( F \), we will consider the following decision problem.

- **\( \text{WinDet}(F) \)**
  
  **Instance:** a agenda \( \Phi \) with an integrity constraint \( \Gamma \), a profile \( J \in J(\Phi, \Gamma)^u \) and a number of subsets \( L_0, \ldots, L_u \subseteq \Phi \) of the agenda, with \( u \geq 0 \).

  **Question:** is there a judgment set \( J^* \in \text{Winner}_{F,\Phi,\Gamma}(J) \) such that \( L_0 \subseteq J^* \) and \( L_i \not\subseteq J^* \) for each \( 1 \leq i \leq u \)?

In addition, we will consider the following search problem.

- **\( \text{FWinDet}(F) \)**
  
  **Input:** an agenda \( \Phi \) with an integrity constraint \( \Gamma \), a profile \( J \in J(\Phi, \Gamma)^u \) and a number of subsets \( L_0, \ldots, L_u \subseteq \Phi \) of the agenda, with \( u \geq 0 \).

  **Output:** a judgment set \( J^* \in \text{Winner}_{F,\Phi,\Gamma}(J) \) such that \( L_0 \subseteq J^* \) and \( L_i \not\subseteq J^* \) for each \( 1 \leq i \leq u \), if it exists.

In these problems, the set \( L_0 \) allows us to specify judgments that the outcome must contain. Moreover, using the sets \( L_1, \ldots, L_u \), we can specify a list of forbidden (combinations of) judgments that the outcome may not contain. This allows us, for instance, to enumerate several outcomes by using previously found outcomes as the sets \( L_1, \ldots, L_u \).

We make several observations about this formalization of the computational task of winner determination for the judgment aggregation procedures. In previous work [13, 21], only decision problems have been studied. We argue that the search problem formulation is a more accurate formalization of the computational task related to the winner determination problem in judgment aggregation. In general, for any judgment aggregation procedure, given a multiagent setting and individual opinions of the agents, one would like to compute a group opinion, which could then be used for various purposes. Given access to an algorithm that performs the computational task captured by the decision problem, one needs a linear number of calls to produce the description of an outcome. However, it turns out that (under some common complexity-theoretic assumptions) any such (deterministic) algorithm invokes an NP oracle only a small (sublinear) number of times in the worst case. This was known for the case of the Kemeny rule [13], and we show that it is also the case for the Slater and Young rules. Therefore, it could be that an algorithm that produces an outcome by simply calling an algorithm for the decision problem a linear number of times exceeds the minimum number of calls to the
NP oracle that are needed. In fact, we show that producing an outcome can be done in deterministic polynomial time with only a logarithmic number of calls to an oracle that produces witnesses for NP problems. Formalizing the task as a search problem in the first place circumvents this issue.

Secondly, these computational problems involve the question whether there exists an outcome that satisfies certain properties, rather than the question whether all outcomes satisfy certain properties (as studied by Lang and Slavkovik [21]). We think the former question is the more natural of these two, as computing outcomes for judgment aggregation procedures is arguably the most central computational task in judgment aggregation.

The third observation concerns the condition that the subsets \( L_1, \ldots, L_u \), given as part of the input to the problems, are not subsets of the judgment set \( J^* \in \text{Winner}_F \). As can be seen in the proofs in the remainder of this section, this condition does not have an effect on the complexity analysis of the problem. However, it does allow us to use this formalization of the problem to devise an algorithm to enumerate outcomes of judgment aggregation procedures. In order to enumerate outcomes, one could use an algorithm that solves \( \text{FWInDet}(F) \) repeatedly, and rule out previously found outcomes by providing them as the subsets \( L_i \) (with \( i > 0 \)).

We obtain the following complexity results for these two computational problems (summarized in Table 1).

**Theorem 1.** For each \( F \in \{ \text{Kemeny}, \text{Slater}, \text{Young} \} \):

a) \( \text{WinDet}(F) \) is \( \text{P}_{\log} \)-complete; and

b) \( \text{FWInDet}(F) \) is \( \text{FP}_{\log, \text{wit}} \)-complete.

**Theorem 2.** For any tie-breaking rule \( A \), it holds that \( \text{WinDet}(\text{RA}_A) \) is in \( \text{P}_{\log} \) and \( \text{FWInDet}(\text{RA}_A) \) is in \( \text{FP}_{\log} \). Moreover, there is some tie-breaking rule \( A \) such that \( \text{WinDet}(\text{RA}_A) \) is \( \text{P}_{\log} \)-complete and \( \text{FWInDet}(\text{RA}_A) \) is \( \text{FP}_{\log} \)-complete.

**Theorem 3.** \( \text{WinDet}(\text{RA}) \) is \( \Sigma_2^0 \)-complete and \( \text{FWInDet}(\text{RA}) \) is \( \text{FP}_{\log} \)-complete.

We would like to point out that the complexity result for \( \text{WinDet}(\text{Kemeny}) \) has been shown before [13]. We give an alternative proof of hardness for this problem. We prove hardness by giving a reduction directly from a canonical complete problem for \( \text{P}_{\log} \), which allows us to extend this hardness result to the case of the search problem \( \text{FWInDet}(\text{Kemeny}) \), which would not have been the case for the search problem \( \text{WinDet}(\text{Kemeny}) \), from which we obtain the known \( \text{P}_{\log} \)-hardness proof for \( \text{WinDet}(\text{Kemeny}) \) from the literature.

We begin by proving Theorem 1 in Sections 3.1–3.3 (Propositions 5–9 and Corollaries 6–10). Then, we prove Theorems 2 and 3 in Section 3.4 (Propositions 11–12 and Corollaries 13–14, and Propositions 15–16 and Corollary 17, respectively).

### 3.1 Membership for Kemeny, Slater, Young

We will show \( \text{FP}_{\log, \text{wit}} \)-membership of the problem \( \text{FWInDet}(F) \), for each \( F \in \{ \text{Kemeny}, \text{Slater}, \text{Young} \} \) by giving a polynomial-time algorithm that uses an FNP witness oracle at most \( O(\log n) \) many times. Since \( \text{P}_{\log} \) coincides with \( \text{P}_{\log, \text{wit}} \), this algorithm can then easily be adapted to show \( \text{FP}_{\log} \)-membership of \( \text{WinDet}(F) \).

The approach that we use to show these membership results is similar to the approach taken in the known \( \text{P}_{\log} \)-membership proof of \( \text{WinDet}(\text{Kemeny}) \) [13]. For each \( F \in \{ \text{Kemeny}, \text{Slater}, \text{Young} \} \), we will consider the following auxiliary problem.

**FSetScore(F)**

**Input:** an agenda \( \Phi \) with an integrity constraint \( \Gamma \), a profile \( J \in \mathcal{J}(\Phi, \Gamma) \), a number of subsets \( L_0, \ldots, L_u \subseteq \Phi \), with \( u \geq 0 \), and a positive integer \( m \) (in unary).

**Output:** a consistent and complete judgment set \( J \) such that \( L_0 \subseteq J \subseteq L_1 \subseteq J \) for each \( i \leq u \).

**Proof (sketch).** We describe a guess-and-check algorithm for \( \text{FSetScore}(F) \). The algorithm gets as input an agenda \( \Phi \) with an integrity constraint \( \Gamma \), a profile \( J \in \mathcal{J}(\Phi, \Gamma) \), a number of subsets \( L_0, \ldots, L_u \subseteq \Phi \), with \( u \geq 0 \), and a positive integer \( m \) in unary. The algorithm guesses a complete judgment set \( J \), and an interpretation \( M \) that witnesses that \( J \) is \( \Gamma \)-consistent. Additionally, if \( F = \text{Young} \), the algorithm guesses \( \ell \leq m \) many judgment sets \( J_1, \ldots, J_\ell \) in \( J \) to remove from \( J \).

Then, the algorithm verifies (1) whether \( M \) satisfies both \( J \) and \( F \); (2) whether \( \text{Score}_F(J, J) \leq m \); (3) whether \( L_0 \subseteq J \); and (4) whether \( L_\ell \subseteq J \) for each \( 1 \leq i \leq u \). If \( F = \text{Young} \), the verification of (2) uses the guessed judgment sets \( J_1, \ldots, J_\ell \). It is straightforward to verify that these checks can be done in polynomial time. If \( J \) satisfies conditions (1)–(4), the algorithm accepts and outputs \( J \); otherwise, the algorithm rejects.

**Proposition 5.** For each \( F \in \{ \text{Kemeny}, \text{Slater}, \text{Young} \} \), \( \text{FWInDet}(F) \in \text{FP}_{\log, \text{wit}} \).

**Proof (sketch).** We describe a polynomial-time algorithm that computes \( \text{FWInDet}(F) \) by using an \( \text{FSetScore}(F) \) oracle. The algorithm firstly computes the minimal number \( m^* \) such that there is some \( \Gamma \)-consistent and complete judgment set \( J \) such that \( \text{Score}_F(J, J) \leq m^* \). For each \( F \in \{ \text{Kemeny}, \text{Slater}, \text{Young} \} \), we know that \( m^* \) is polynomially bounded by the size of the encoding of \( J \). Therefore, we can employ a binary search strategy to compute \( m^* \) by asking \( O(\log |J|) \) many oracle queries of the form \( \langle \Phi, \Gamma, J, \emptyset, m \rangle \). Then, the algorithm returns the answer to the oracle query \( \langle \Phi, \Gamma, J, L_0, \ldots, L_\ell, m^* \rangle \).

**Corollary 6.** For each \( F \in \{ \text{Kemeny}, \text{Slater}, \text{Young} \} \), \( \text{WinDet}(F) \in \text{P}_{\log} \).

### 3.2 Hardness for Kemeny and Slater

**Proposition 7.** The problems \( \text{FWInDet}(\text{Kemeny}) \) and \( \text{FWInDet}(\text{Slater}) \) are \( \text{FP}_{\log, \text{wit}} \)-hard, even with the restriction that \( \Gamma = T \), and \( u = 0 \).

**Proof (sketch).** We give a polynomial-time Levin reduction from \( \text{X-MAX-MODEL} \). The same reduction works for both \( \text{FWInDet}(\text{Kemeny}) \) and \( \text{FWInDet}(\text{Slater}) \). Let \((\varphi, X)\) be an input to \( \text{X-MAX-MODEL} \) with \( X = \{ x_1, \ldots, x_n \} \subseteq \text{Var}(\varphi) \). We construct an agenda \( \Phi \), an integrity constraint \( \Gamma = T \), an integer \( u = 0 \), a subset \( L_0 \subseteq \Phi \), and a profile \( J \in \mathcal{J}(\Phi, \Gamma)^n \) as follows. We pick some \( w \)}
such that \( w > v^2 \), e.g., \( w = 2v^2 \). For each \( 1 \leq i \leq v \) and \( 1 \leq j \leq w \), we introduce a fresh variable \( z_{j,i} \). Moreover, we introduce a fresh variable \( y \). Then, we construct a propositional formula \( \varphi' \) that is true if and only if one of the following conditions holds: either (i) for some \( 1 \leq i \leq v \), \( x_i \) is set to false, \( y \) is set to false, and all \( z_{j,i} \) for \( 1 \leq j \leq w \) are set to true, or (ii) all \( z_{j,i} \) are set to false, \( y \) is set to true, and the assignment to the variables \( x_i \) satisfies \( \varphi \). It is straightforward to construct in polynomial time some formula \( \varphi' \) that satisfies this property. Then we introduce \( w \) many syntactic copies \( \varphi'_j \) of \( \varphi' \). We now define the agenda \( \Phi \) by letting \( \{ \Phi \} = \{ y \} \cup \{ x_1, \ldots, x_v \} \cup \{ z_{j,i} : 1 \leq j \leq w, 1 \leq i \leq n \} \cup \{ \varphi'_1, \ldots, \varphi'_w \} \). Next, we construct the profile \( J = (J_1, \ldots, J_w) \) as shown in Figure 1. Moreover, we let \( L_0 = \{ y \} \).

\[
\begin{array}{c|cccc|c}
 J & J_1 & J_2 & \ldots & J_{v-1} & J_v & m(J) \\
\hline
 y & 0 & 0 & 0 & \ldots & 0 & 0 \\
 x_1 & 0 & 1 & 1 & \ldots & 1 & 1 \\
 x_2 & 1 & 0 & 1 & \ldots & 1 & 1 \\
 \vdots & \vdots & \vdots & \ddots & \vdots & \vdots \\
 x_{v-1} & 1 & \ldots & 1 & 0 & 1 & 1 \\
 x_v & 1 & \ldots & 1 & 1 & 0 & 1 \\
 z_{1,1} & 1 & 0 & 0 & \ldots & 0 & 0 \\
 z_{1,2} & 0 & 1 & 0 & \ldots & 0 & 0 \\
 \vdots & \vdots & \vdots & \ddots & \vdots & \vdots \\
 z_{w,v-1} & 0 & \ldots & 0 & 1 & 0 & 0 \\
 z_{w,v} & 0 & \ldots & 0 & 0 & 1 & 0 \\
 \varphi'_1 & 1 & 1 & \ldots & 1 & 1 \\
 \varphi'_2 & 1 & 1 & \ldots & 1 & 1 \\
 \vdots & \vdots & \vdots & \ddots & \vdots & \vdots \\
 \varphi'_{w-1} & 1 & \ldots & 1 & 1 & 1 \\
 \varphi'_w & 1 & \ldots & 1 & 1 & 1 \\
\end{array}
\]

Figure 1: construction of the profile in the proof of Proposition 7.

What remains is to specify a polynomial-time computable function \( g \) that takes some \( J^* \in \text{Winner}_{\psi, \Gamma}(J) \) with \( L_0 \subseteq J^* \), if it exists, and that produces a model \( M \) such that \( X\text{-Max-Model}(\varphi, X, M) \). Let \( J^* \in \text{Winner}_{\psi, \Gamma}(J) \) such that \( L_0 \in J^* \). Then the function \( g \) outputs \( J^* \cap \{ x_1, \ldots, x_v \} \).

The intuition behind this reduction is the following. The agenda and the profile are constructed in such a way that it is ‘cheaper’ to agree with the overall profile (respectively, with the majority outcome) on the formulas \( \varphi'_j \) and the formulas \( z_{j,i} \), if possible. If \( \varphi \) is satisfiable, the cheapest way to agree with the overall profile (respectively, with the majority outcome) on the above formulas (in a consistent judgment set) is to set the variables \( x_i \) to a maximal assignment \( \alpha \) to the variables in \( X \) that is extendable to a satisfying assignment to \( \varphi \), and to set \( y \) to true. If \( \varphi \) is unsatisfiable, all consistent judgment sets that disagree with the majority outcome on a minimal number of formulas (e.g., setting \( z_{j,i} \), true for some \( 1 \leq i \leq v \) and \( 1 \leq j \leq w \) do not contain \( y \) (and thus are not a superset of \( L_0 \))

**Corollary 8.** The problems \( \text{WinDet}(\text{Kemeny}) \) and \( \text{WinDet}(\text{Slater}) \) are \( \text{P}^{\text{NP}[\log]} \)-hard, even with the restriction that \( \Gamma = \top \), and \( u = 0 \).

**Proof (sketch).** The reduction in the proof of Proposition 7 can be adapted to a many-one reduction from \( \text{WinDet}(\text{Kemeny}) \) (respectively, from \( \text{WinDet}(\text{Slater}) \)) to \( X\text{-Max-Model-Parity} \), by letting \( L_0 = \{ y, x_0 \} \), where \( x_0 \) is given in the original input for the problem \( X\text{-Max-Model-Parity} \).

### 3.3 Hardness for Young

**Proposition 9.** \( F\text{WinDet}(\text{Young}) \) is \( \text{FP}^{\text{NP}[\log, \text{wit}]} \)-hard, even with the restriction that \( \Gamma = \top \), and \( u = 0 \).

**Proof (sketch).** We give a polynomial-time Levin reduction from \( X\text{-Max-Model} \). Let \( \varphi \) be an input to \( X\text{-Max-Model} \) with \( \text{Var}(\varphi) = \{ x_1, \ldots, x_v \} \). We may assume that any satisfying assignment (if there exists any) sets at least one variable in \( X \) to true, and that setting all variables in \( X \) to true does not satisfy \( \varphi \). We construct an agenda \( \Phi \), an integrity constraint \( \Gamma = \top \), an integer \( u = 0 \), a subset \( L_0 \subseteq \Phi \), and a profile \( J \in J(\Phi, \Gamma)^n \) as follows. We let \( w = 2v + 1 \). Then, we construct a propositional formula \( \varphi' \) that is satisfiable if and only if one of the following conditions holds: either (i) for some \( 1 \leq j \leq w \), \( y_i \) is set to true, or (ii) all \( y_i \) are set to false and the assignment to the variables \( x_i \) reduces \( \varphi \) to a satisfiable formula. It is straightforward to construct in polynomial time some formula \( \varphi' \) that satisfies this property. We now define the agenda \( \Phi \) by letting \( \{ \Phi \} = \{ x_1, \ldots, x_v \} \cup \{ y_1 : 1 \leq j \leq w \} \cup \{ \varphi' \} \). Next, we construct the profile \( J = (J_1, \ldots, J_w) \) as shown in Figure 2. Moreover, we let \( L_0 = \{ \neg y_1, \ldots, \neg y_w \} \).

What remains is to specify a polynomial-time computable function \( g \) that takes some \( J^* \in \text{Winner}_{\text{Young}, \psi, \Gamma}(J) \) with \( L_0 \subseteq J^* \), if it exists, and that produces a model \( M \) such that \( X\text{-Max-Model}(\varphi, M) \). Let \( J^* \in \text{Winner}_{\text{Young}, \psi, \Gamma}(J) \). Then the function \( g \) outputs \( J^* \cap \{ x_1, \ldots, x_v \} \).

The intuition behind this reduction is the following. Since the profile is unanimous on \( \varphi' \), any winner must include \( \varphi' \). Satisfying \( \varphi' \) can be done by either setting some \( y_i \) to true, or by satisfying the original formula \( \varphi \). In the case that \( \varphi \) is satisfiable, the only subprofiles that have a consistent majority outcome are subprofiles consisting of a single judgment set \( J \), for which all holds that \( L_0 \not\subseteq J \).

**Corollary 10.** \( \text{WinDet}(\text{Young}) \) is \( \text{P}^{\text{NP}[\log]} \)-hard, even with the restriction that \( \Gamma = \top \), and \( u = 0 \).

**Proof (sketch).** The reduction in the proof of Proposition 9 can be adapted to a many-one reduction from \( \text{WinDet}(\text{Young}) \) to \( X\text{-Max-Model-Parity} \), by letting \( L_0 = \{ x_0, x_1 \} \), where \( x_0 \) is given in the original input for the problem \( X\text{-Max-Model-Parity} \).
3.4 Completeness for Tideman

**Proposition 11.** Let \( <^A \) be some tie-breaking rule. Then \( \WinDet(RA_A) \) is in \( \text{FP}^{NP} \).

**Proof (sketch).** We describe a deterministic polynomial-time algorithm with access to a (yes-no) SAT oracle that solves \( \WinDet(RA_A) \). The algorithm takes as input an agenda \( \Phi \), an integrity constraint \( \Gamma \), a profile \( J \in \mathcal{T}(\Phi, \Gamma)^u \), and a number of subsets \( L_0, \ldots, L_u \) of the agenda with \( u \geq 0 \). In polynomial time, the algorithm computes \( \leq_J \), and computes the Ranked-Agenda order \( <^J_L \), using the tie-breaking rule \( <^A \). Let \( \Phi = \{ \varphi_1, \ldots, \varphi_m \} \), where \( \varphi_i >^A \cdots >^A \varphi_2m \). We compute \( \text{RA}(<^J_L, \Phi, \Gamma) = J_{2m} \) by iteratively computing \( J_i \), for all \( 0 \leq i \leq 2m \). Let \( J_0 = \emptyset \). Given \( J_i \), we compute \( J_{i+1} \) as follows, by querying the SAT oracle. We query the SAT solver whether \( \psi_{i+1} = \bigwedge_{\varphi \in \psi_{i+1}} \varphi \wedge \wedge_{\varphi \in \psi_{i+1}} \varphi \) is \( \Gamma \)-consistent. If \( \psi_{i+1} \) is \( \Gamma \)-consistent, we let \( J_{i+1} = J_i \cup \{ \psi_{i+1} \} \); otherwise, we let \( J_{i+1} = J_i \). Clearly, this requires \( 2m = |\Phi| \) calls to the SAT oracle. Then, let \( J^* = J_{2m} = \text{RA}(<^J_L, \Phi, \Gamma) \). Finally, the algorithm verifies whether \( \forall L_0 \subseteq J^* \) and \( L_i \subseteq J^* \) for all \( 1 \leq i \leq u \). If these checks do not succeed, the algorithm rejects the input; otherwise, it returns \( J^* \). \( \square \)

**Proposition 12.** \( \WinDet(RA_A) \) is \( \text{FP}^{NP} \)-hard for some (fixed) tie-breaking rule \( A \), even with the restriction that \( \Gamma = \top \), and \( u = 0 \).

**Proof (sketch).** We give a polynomial-time Levin reduction from \( \text{FLEX-MAX-MODEL} \). Let \( \varphi \) be an input for \( \text{FLEX-MAX-MODEL} \), with \( \text{Var}(\varphi) = \{ x_1, \ldots, x_v \} \). We construct an agenda \( \Phi \), an integrity constraint \( \Gamma = \top \), an integer \( u = 0 \), a linear order \( <^L \subseteq \Phi \times \Phi \), a profile \( J \in \mathcal{T}(\Phi, \Gamma)^u \), and a subset \( L_0 \subseteq \Phi \), as follows. We pick some even \( w \) such that \( w > v \), e.g., \( w = 2v \). We introduce auxiliary variables \( y_1, \ldots, y_w \). Then, we construct a propositional formula \( \varphi' \) that is true if and only if one of the following conditions holds: (i) for some \( 1 \leq j \leq w \), \( y_j \) is set to true, or (ii) the assignment to the variables \( x_i \) satisfies \( \varphi \). We define the agenda \( \Phi \) by letting \( \Phi = \{ \varphi' \} \cup \{ y_1, \ldots, y_w \} \cup \{ x_1, \ldots, x_v \} \cup \{ \} \). Then, we define the total order \( <^A \) by letting \( <^A \gamma <^A \cdots <^A \gamma x \cdots <^A \gamma x <^A x \cdots <^A x <^A \gamma \). Next, we construct the profile \( J = (J_1, \ldots, J_w) \) as shown in Figure 3. Note that the Ranked-Agenda order \( <^J_L \) coincides with the order \( <^A \). Finally, we let \( L_0 = \{ \gamma \} \).

![Figure 2: construction of the profile in the proof of Proposition 9.](image1)

![Figure 3: construction of the profile in the proof of Proposition 12.](image2)

What remains is to specify a polynomial-time computable function \( g \) that takes some \( J^* \in \text{Winner}_{\text{RA}A_\varphi}(J) \) with \( L_0 \subseteq J^* \), and that produces a model \( M \) such that \( \text{FLEX-MAX-MODEL}(\varphi, M) \). The function \( g \) outputs \( J^* \cap \{ x_1, \ldots, x_v \} \).

Note that for each instance \( \varphi \) of \( \text{FLEX-MAX-MODEL} \), this reduction produces a different agenda \( \Phi \). Therefore, there is a single tie-breaking rule \( A \) that produces the correct order \( <^A \) for each agenda \( \Phi \) that is generated by this reduction (and moreover, this order \( <^A \) clearly is computable in polynomial time, given \( \varphi \)).

The intuition behind this reduction is the following. If \( \varphi \) is satisfiable, the Ranked-Agenda winner (for the tie-breaking rule \( A \)) will contain \( \varphi' \) and \( \neg y_1, \ldots, \neg y_w \), because these are consistent (and are the first formulas in the order \( <^J_L \)), and if \( \varphi \) is unsatisfiable, the winner will not contain \( \neg y_w \), because \( \varphi' \) and \( \neg y_1, \ldots, \neg y_{w-1} \) are not consistent with \( \neg y_w \). Moreover, if \( \varphi \) is satisfiable, the Ranked-Agenda order \( <^J_L \) forces the winner to contain the lexicographically maximal model of \( \varphi \).

**Corollary 13.** Let \( <^A \) be some tie-breaking rule. Then \( \WinDet(RA_A) \) is in \( \text{P}^{NP} \).

**Corollary 14.** \( \WinDet(RA_A) \) is \( \text{P}^{NP} \)-hard for some (fixed) tie-breaking rule \( A \), even with the restriction that \( \Gamma = \top \), and \( u = 0 \).

**Proof (sketch).** The reduction in the proof of Proposition 12 can be adapted to a many-one reduction from
Proposition 15. \textsc{FWinDet}(RA) is in $\text{F} \Sigma_2^p$.

Proof (sketch). We describe a nondeterministic polynomial-time algorithm with access to a (yes-no) SAT oracle that solves \textsc{FWinDet}(RA). The algorithm takes as input an agenda $\Phi$, an integrity constraint $\Gamma$, a profile $J \in J(\Phi,\Gamma)^n$, and a number of subsets $L_0, \ldots, L_u \subseteq \Phi$ of the agenda. It is straightforward to compute $\leq_J$ in polynomial time. Then, the algorithm guesses some total order $<_J^t \subseteq \Phi \times \Phi$. It is straightforward to verify in polynomial time that $<_J^t$ extends the partial order $\leq_J$. (If $<_J^t$ does not extend $\leq_J$, the algorithm rejects.) Then, the algorithm proceeds similarly to the algorithm described in the proof of Proposition 11 to compute $J' = RA(<_J^t, \Phi, \Gamma)$. Also, the algorithm verifies whether $L_0 \subseteq J'$ and $L_i \not\subseteq J'$ for all $1 \leq i \leq u$. If these checks do not succeed, the algorithm rejects; otherwise, it returns $J'$. □

Proposition 16. \textsc{FWinDet}(RA) is $\text{F} \Sigma_2^p$-hard, even with the restriction that $\Gamma = \top$, and $u = 0$.

Proof (sketch). We give a polynomial-time Levin reduction from \textsc{FQSAT}_\text{FQ}. Let $\varphi = 3X.\forall Y.\psi$ be an instance of \textsc{FQSAT}_\text{FQ}, where $X = \{x_1, \ldots, x_\ell\}$. We may assume without loss of generality that $\psi$ is not a tautology. We construct an agenda $\Phi$, an integrity constraint $\Gamma = \top$, an integer $u = 0$, a profile $J \in J(\Phi,\Gamma)^n$, and a subset $L_0 \subseteq \Phi$ of the agenda, as follows. We pick some $w$ such that $w > v$, e.g., $w = 2v$. We introduce auxiliary variables $y_1, \ldots, y_w$. Also, we introduce auxiliary variables $z_i$ for each $l \in \{x_1, \ldots, x_\ell\}$. Then we construct a propositional formula $\varphi$ that is true if and only if one of the following two conditions holds: either (i) some $y_i$ is true, or (ii) for each $l \in \{x_1, \ldots, x_\ell\}$ it holds that $l$ is true if and only if $z_i$ is true. Moreover, we construct a propositional formula $\psi$ that is true if and only if both of the following two conditions holds: both (i) no $y_i$ with $1 \leq i \leq 3$ is true, and (ii) either (ii.a) some $y_i$ is true, or (ii.b) the formula $\neg \psi$ is satisfiable. It is straightforward to construct these formulas $\varphi$ and $\psi$ in polynomial time. We now define the agenda $\Phi$ by letting $[\Phi] = \{\chi, \psi\} \cup \{z_1, z_{\ell-1} : 1 \leq i \leq \ell\} \cup \{y_1, \ldots, y_w\}$. Next, we construct the profile $J = (J_1, \ldots, J_u)$ as shown in Figure 4. Note that $m(J, \chi) = w$; $m(J, \psi') = w - 3$; for each $1 \leq i \leq v$, $m(J, x_i) = m(J, z_{x_i}) = w - 1$; for each $1 \leq j \leq w$, $m(J, y_j) = w - 1$. Therefore $\{\chi\} \leq_J \{\neg y_j : 1 \leq j \leq w\} \leq_J \{z_i : 1 \leq i \leq \ell\} \leq_J \psi$. Moreover, we let $L_0 = \{\neg \psi\}$.

What remains is to specify a polynomial-time computable function $g$ that takes some $J' \in \text{Winner}_{RA,\Phi,\Gamma}(J)$ with $L_0 \subseteq J'$, and that produces a model $M$ such that $\textsc{FQSAT}_\text{FQ}(\varphi, M)$. Let $J' \in \text{Winner}_{RA,\Phi,\Gamma}(J)$. The function $g$ outputs $\{x_i : 1 \leq i \leq v, z_i \in J\}$. □

Figure 4: construction of the profile in the proof of Proposition 16.

The intuition behind this reduction is the following. The profile is constructed in such a way that every Ranked-Agenda winner contains $\chi$. If there is an assignment $\alpha$ to the variables $\{x_1, \ldots, x_\ell\}$ such that $\forall Y.\psi[\alpha]$ is true, then there exists a winner containing all literals $\neg y_j$, the variables $z_i$ for those literals that are satisfied by $\alpha$, and the formula $\neg \psi'$, because any set containing $\chi$ the literals $\neg y_j$, and those variables $z_i$ would be inconsistent with $\psi'$, since such a set implies $\psi$ (and thus forces $\neg \psi$ to be unsatisfiable). Moreover, any winner containing $\neg \psi'$ is of such a form, and thus corresponds to an assignment $\alpha$ with the property that $\forall Y.\psi[\alpha]$ is true. □

Corollary 17. \textsc{WinDet}(RA) is $\Sigma_2^p$-complete. Hardness holds even with the restriction that $\Gamma = \top$, and $u = 0$.

4. CONCLUSION

We studied the computational complexity of the winner determination problem for (the judgment aggregation analogues of) the Kemeny rule, the Slater rule and the Young rule, as well as two variants of the Ranked-Agenda rule. These computational tasks we formalized in the form of a search problem and a decision problem. We showed that the complexity for the Kemeny, Slater and Young rules lies at the $\Theta_2^p$-level of the PH, and we showed that the complexity of the Ranked-Agenda rule, lies at the $\Delta_2^p$-level of the PH, in case of a fixed tie-breaking rule, and lies at the $\Sigma_2^p$-level of the PH in case ties can be broken in arbitrary ways.

Future research should include investigating the computational complexity of the winner determination problem for other complete and consistent judgment aggregation procedures. One example of such a procedure is one based on the distance measure introduced by Duddy and Piggins [11]. Another direction for further research is to investigate the contribution of various aspects of the problem to its computational complexity, by analyzing the problems using the framework of parameterized complexity theory.

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