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REDUCIBILITIES ON TALLY AND SPARSE SETS (*)

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Abstract. — Classes of sets that are inter-reducible to tally sets by polynomial-time computable reducibilities are studied.

Résumé. — On considère les classes d'ensembles qui sont inter-réductibles vis-à-vis des ensembles de comptage (tally sets) par réductibilités calculables en polynôme-temps.

1. INTRODUCTION

The notions of tally set and sparse set represent the best known examples of sets with "small information content". As such, sparse sets and tally sets have played important roles in the investigation of the structure of complexity classes and of polynomial time-bounded reducibilities. For example, the class of sets that have polynomial-size circuits is the class of sets that are Turing reducible in polynomial time to sparse sets; the class of sets with small generalized Kolmogorov complexity is precisely the class of sets that are polynomial-time isomorphic to tally sets (Allender and Rubinstein [1]); and the class of sets with self-producible circuits (as defined by Ko [8]) is precisely the class of sets that are Turing equivalent in polynomial-time to tally sets (Balcázar and Book [4]). These results, as well as those of Book and Ko [6],

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are considered to be steps towards understanding how different resource-bounded reducibilities can (and cannot) be used to retrieve information encoded in sets with small information content.

Let \( \text{SPARSE} \) denote the set of all sparse sets and let \( \text{TALLY} \) denote the set of all tally sets. For any of the standard (i.e., many-one \( \leq_m \), \( k \)-truth-table \( \leq_k \), bounded truth-table \( \leq_{bt} \), truth-table \( \leq_t \), Turing \( \leq_T \)) reducibilities \( R \) and any class \( C \) of sets, let \( P^R_R(C) = \{ A \mid \text{there exists } C \in C \text{ such that } A \leq^R_R C \} \) and let \( E^R_R(C) = \bigcup_{C \in C} \{ A \mid A \leq^R_R C \text{ and } C \leq^R_R A \} \); \( P^R_R(C) \) is the reduction class of \( C \) under \( \leq^R_R \). [If \( C \) consists of a single set, \( C = \{ C \} \), then we write \( P^R_R(Q) \) for \( P^R_R(C) \) and \( E^R_R(Q) \) for \( E^R_R(C) \).]

Book and Ko [6] observed that the class of sets with polynomial size circuits, denoted \( P/\text{poly} \), has the property that

\[
P/\text{poly} = P_T (\text{SPARSE}) = P^tt (\text{SPARSE}) = P^tt (\text{TALLY}) = P_T (\text{TALLY}).
\]

In addition, \( P_{bt} (\text{SPARSE}) \neq P/\text{poly} \) and for every \( k > 0 \),

\[
P_{k-\text{tt}} (\text{SPARSE}) \neq P^{(k+1)-\text{tt}} (\text{SPARSE}).
\]

Thus, the class \( P/\text{poly} \) can be decomposed into an infinite hierarchy of classes based on the number of queries made to sparse oracles. In contrast, Book and Ko showed that

\[
P_m (\text{TALLY}) = P_{bt} (\text{TALLY}) \quad \text{and} \quad P_{bt} (\text{TALLY}) \neq P^tt (\text{TALLY})
\]

so that when considering the number of queries made to tally oracles, there are only two classes, \( P/\text{poly} \) and \( P_m (\text{TALLY}) \).

Since for any tally set \( T \), \( P^tt (T) = P_T (T) \), neither an individual tally set nor the union of the reduction classes of all tally sets can be used to distinguish between polynomial time truth-table reducibility and polynomial time Turing reducibility. But the situation is different when the classes of sets interreducible to tally sets under these reducibilities are considered: we show in Theorem 3.5 that the resulting classes are not the same, more specifically, that \( E^p^tt (\text{TALLY}) \neq E^p_T (\text{TALLY}) \); thus, there are sets with self-producible circuits that are not truth-table equivalent in polynomial time to tally sets. Hence, it is possible to hide information about a tally set in a non-tally set in such a way that it is retrievable by Turing reductions to that non-tally set but it is no retrievable by truth-table reductions to the same set.

In addition, we consider strong nondeterministic polynomial-time Turing reducibility as defined by Long [11] and we show in Theorem 3.3 that
\[ E_{SN}^P(\text{TALLY}) \neq E_{T}^P(\text{TALLY}) \]. Thus, we have the proper inclusions

\[ E_{n}^P(\text{TALLY}) \subseteq E_{T}^P(\text{TALLY}) \subseteq E_{SN}^P(\text{TALLY}). \]

Whether the corresponding inclusions

\[ E_{n}^P(\text{SPARSE}) \subseteq E_{T}^P(\text{SPARSE}) \subseteq E_{SN}^P(\text{SPARSE}) \]

are proper is an open question.

Based on the fact that for every \( k \), \( P_{k+1}(\text{SPARSE}) \neq P_{n}(\text{SPARSE}) \), we observe that \( E_{n}^P(\text{SPARSE}) \neq E_{n}^P(\text{SPARSE}) \). But the analogous situation with TALLY is unresolved. It can be shown that for each \( k > 0 \), there exists a tally set \( T \) such that \( E_{k+1}^P(\text{TALLY}) \neq E_{k}^P(\text{TALLY}) \). But for each \( k \), the question \( E_{k+1}^P(\text{TALLY}) = ? E_{k}^P(\text{TALLY}) \) is open, as are the questions \( E_{m}^P(\text{TALLY}) = ? E_{m}^P(\text{TALLY}) \) and \( E_{n}^P(\text{TALLY}) = ? E_{n}^P(\text{TALLY}) \).

Allender and Watanabe \[2\] have shown each of these three questions to be equivalent to the question of whether every honest polynomial-time computable function \( f : \Sigma^* \rightarrow \{0\}^* \) is weakly invertible. Thus, there is reason to believe that these questions will be difficult to resolve.

2. PRELIMINARIES

In this section we review some definitions and establish notation additional to that given in the Introduction.

Throughout this paper we will consider the alphabet \( \Sigma = \{0, 1\} \). The length of a string \( x \) will be denoted by \( |x| \). The cardinality of a set \( S \) will be denoted by \( |S| \). For a set \( S \) and an integer \( n \), \( S^n = \{ x \in S \mid |x| = n \} \) and \( S^\leq n = \{ x \in S \mid |x| \leq n \} \). For a set \( S \), \( \chi_S \) denotes the characteristic function of \( S \), and if \( S \subseteq \Sigma^* \), then \( \tilde{S} = \Sigma^* - S \).

A set \( S \) is sparse if there is a polynomial \( q \) such that for all \( n \), \( |S^\leq n| \leq q(n) \). Let SPARSE denote the class of all sparse sets. A tally set is any subset of \( \{0\}^* \). Let TALLY denote the class of all tally sets.

We will assume the existence of a pairing function \( \langle , \rangle : \Sigma^* \times \Sigma^* \rightarrow \Sigma^* \) with the properties that function and its inverses are computable in polynomial time and that when restricted to \( \{0\}^* \times \{0\}^* \) yields a string in \( \{0\}^* \).

For an oracle machine \( M \), \( L(M, A) \) denotes the set of strings accepted by \( M \) relative to oracle set \( A \), and \( L(M) \) denotes the set of strings accepted by \( M \) when no oracle queries are allowed by \( M \).
We assume that the reader is familiar with the well-studied polynomial-
time computable reducibilities referred to in the Introduction, with the 
corresponding reducibilities that are computed nondeterministically (i.e., 
\(\leq_{NP}^{T}\), etc.), with the complexity classes \(P, NP,\) and \(PSPACE\) and their relativizations, and with the basic properties of these notions. An appropriate (and sufficiently comprehensive) reference for these topics is the book by Balcázar, Diaz, and Gabarró [5].

Long [11] studied the notion of "strong" nondeterministic polynomial-time 
reducibilities; \(\leq_{SN}^{T}\) denotes strong nondeterministic Turing reducibility. For 
our purposes it is sufficient to use the following characterization of 
\(\leq_{SN}^{T}: A \leq_{SN}^{T} B\) if \(A \leq_{NP}^{T} B\) and \(\overline{A} \leq_{NP}^{T} B\). (See [5].) Consistent with the notation 
established in Section 1, we let

\[ E_{NP}^{T} (\text{SPARSE}) = \bigcup_{S \in \text{SPARSE}} \{ A \mid A \leq_{NP}^{T} S \text{ and } S \leq_{NP}^{T} A \}, \\
E_{SN}^{T} (\text{TALLY}) = \bigcup_{T \in \text{TALLY}} \{ A \mid A \leq_{SN}^{T} T \text{ and } T \leq_{SN}^{T} A \}, \\
E_{T}^{SN} (\text{SPARSE}) = \bigcup_{S \in \text{SPARSE}} \{ A \mid A \leq_{SN}^{T} S \text{ and } S \leq_{SN}^{T} A \}. \]

The reader should be cautioned. It is known (see [3] or [5]) that the 
reducibility \(\leq_{NP}^{T}\) is not transitive so that "\(A \leq_{NP}^{T} S\) and \(S \leq_{NP}^{T} A\)" is not an 
equivalence relation. Thus, for a set \(A, E_{NP}^{T} (A)\) is not what might be called 
the "nondeterministic polynomial-time Turing degree of \(A\)". Similarly, it is 
known (see [10] or [5]) that for each \(k > 0, \leq_{k-\text{ut}}^{P}\) is not transitive, but that 
\(\leq_{ut}^{P}\) and \(\leq_{T}^{P}\) are transitive and so "\(A \leq_{T}^{P} S\) and \(S \leq_{T}^{P} A\)" and "\(A \leq_{T}^{P} S\) and \(S \leq_{T}^{P} A\)" are equivalence relations. In addition, it is known (see [11] or [5]) that \(\leq_{SN}^{T}\) is transitive.

3. MAIN RESULTS

In this section we establish the main results of this paper. There are various 
relationships between the various classes defined from the class TALLY of 
at all tally sets and from the class SPARSE of all sparse sets that follow 
naturally from the definitions. The principal new results are the inequalities 
\(E_{T}^{P} (\text{TALLY}) \neq E_{SN}^{P} (\text{TALLY})\) and \(E_{T}^{P} (\text{SPARSE}) \neq E_{T}^{P} (\text{TALLY})\), established in 
Theorems 3.3 and 3.5 respectively. Other results follow from relationships 
established elsewhere but are given here for the sake of completeness.

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We begin by establishing an equality.

**Theorem 3.1:** $E^\text{NP}_T(\text{SPARSE}) = E^\text{NP}_T(\text{TALLY})$.

**Proof:** Since $\text{TALLY} \subseteq \text{SPARSE}$, it follows that $E^\text{NP}_T(\text{TALLY}) \subseteq E^\text{NP}_T(\text{SPARSE})$.

To see the converse, recall that for every $S \in \text{SPARSE}$ there exists $T_S \in \text{TALLY}$ such that $S \leq_T^\text{NP} T_S$ [7]; the proof of this fact given by Schöning (Theorem 4.6 of [13]) shows that $T_S \leq_T^\text{NP} S$ and the reduction of $T_S$ to $S$ can be implemented by a nondeterministic oracle machine in which acceptance corresponds to computations where all of the oracle queries yield the answer "yes"; hence, the reduction $T_S \leq_T^\text{NP} S$ is a conjunctive reduction. For any sparse set $S$, if $A \in E^\text{NP}_T(S)$, then $A \leq_T^\text{NP} T_S$ since $A \leq_T^\text{NP} S$ and $S \leq_T^\text{NP} T_S$, the latter reduction being carried out deterministically; also, $T_S \leq_T^\text{NP} A$ since $T_S \leq_T^\text{NP} S$ by means of a conjunctive reduction and $S \leq_T^\text{NP} A$ (recall that $\leq_T^\text{NP}$ is not transitive so the use of some modification of $\leq_T^\text{NP}$ such as being conjunctive is necessary). Thus, for any sparse set $S$, if $A \in E^\text{NP}_T(S)$, then $A \in E^\text{NP}_T(T_S)$. Hence, $E^\text{NP}_T(\text{SPARSE}) \subseteq E^\text{NP}_T(\text{TALLY})$.

Consider strong nondeterministic polynomial time reducibilities studied by Long [11].

**Theorem 3.2:**

$E^\text{SN}_T(\text{TALLY}) \neq E^\text{NP}_T(\text{TALLY})$, $E^\text{SN}_T(\text{TALLY}) \neq E^\text{SN}_T(\text{SPARSE})$, and for each reducibility $R \in \{ \leq_m, \leq_k, \leq_{bt}, \leq_{it}, \leq_T \}$, $E^R_T(\text{TALLY}) \neq E^R_T(\text{SPARSE})$.

**Proof:** Since $E^\text{NP}_T(\text{TALLY}) = E^\text{NP}_T(\text{SPARSE})$ by Theorem 3.1, we have

$\text{SPARSE} \subseteq E^R_T(\text{SPARSE}) \subseteq E^\text{NP}_T(\text{TALLY})$.

Long [12] has shown that there exists $S \in \text{SPARSE}$ such that for all $T \in \text{TALLY}$, $S \notin E^R_T$. Thus, SPARSE is not included in $E^\text{SN}_T(\text{TALLY})$ so SPARSE is not included in $E^R_T(\text{TALLY})$ and $E^R_T(\text{TALLY}) \neq E^R_T(\text{SPARSE})$ for any $R$. It follows that $E^\text{SN}_T(\text{TALLY}) \neq E^\text{SN}_T(\text{TALLY})$ since

$\text{SPARSE} \subseteq E^\text{NP}_T(\text{SPARSE}) = E^\text{NP}_T(\text{TALLY})$. 

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The problem $E^S_T(SPARSE) = ? E^N_T(SPARSE)$ remains open.

Consider the class $E^P_T(TALLY)$. The inequality

$$E^P_T(TALLY) \neq E^P_T(SPARSE)$$

is established in Theorem 3.2 above. We show that $E^P_T(TALLY)$ is properly included in $E^S_T(SPARSE)$; the question of whether $E^P_T(SPARSE)$ is properly included in $E^S_T(SPARSE)$ remains open.

**Theorem 3.3:** $E^P_T(TALLY) \neq E^S_T(TALLY)$.

**Proof:** A set $S \subseteq \{0, 1\}^*$ will be called *special* if for all $n \geq 1$, $S$ has exactly one element of length $n$. Thus, every special set is sparse. For any special set $S$, let $T(S) = \{<0^n, 0^i>|1 \leq i \leq n$ and the $i$-th bit of the unique element of length $n$ in $S$ is 0\}$; recall that we assume that the pairing function $< , >$ used here satisfies the condition that $x, y \in \{0\}^*$ implies $<x, y> \in \{0\}^*$, so that $T(S)$ is a tally set.

It is clear that $S \subseteq E^P_T(S)$ so that $S \subseteq E^S_T(S)$. Notice that $T(S) \subseteq E^P_T(S)$ since on input a string $<0^n, 0^i>$ a nondeterministic machine can determine whether $1 \leq i \leq n$, can guess a string $y$ of length $n$ whose $i$-th bit is 0, and determine (by querying the oracle) whether $y \in S$; if all conditions are satisfied, then the machine can accept $<0^n, 0^i>$. Similarly, notice that $\overline{T}(S) \subseteq E^P_T(S)$ since on input a string $<0^n, 0^i>$ a nondeterministic machine can determine whether $1 \leq i \leq n$, can guess a string $y$ of length $n$ whose $i$-th bit is 1, and determine (by querying the oracle) whether $y \in S$; if all conditions are satisfied, then the machine can accept $<0^n, 0^i>$. Thus $T(S) \subseteq E^S_T(S)$. Hence, $S \subseteq E^S_T(T(S))$ and $T(S) \subseteq E^S_T(S)$ so that $S \in E^S_T(T(S)) \subseteq E^S_T(TALLY)$.

We claim that there is a special set that is not in $E^P_T(TALLY)$ so that $E^P_T(TALLY) \neq E^S_T(TALLY)$.

Let $S$ be a special set with the property that for each $n$ of the form

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the unique string $s_n$ of length $n$ in $S$ is random in the sense that $s_n$ has Kolmogorov complexity greater than $n/2$; let $s_n = 0^n$ for other $n$. The fact that each string $s_n$ has Kolmogorov complexity greater than $n/2$ may be interpreted as saying that $S$ is not self-reducible.

Suppose $S \in E^P_T(TALLY)$. Let $T$ be a tally set such that $S \subseteq T$ and $T \subseteq S$. Let $M$ be a deterministic oracle machine that witnesses $T \subseteq S$, and let $q(n)$ be a polynomial that bounds $M$'s running time. For each $n$, every tally string in $T$ of length at most $q(n)$ can be recognized by $M$ relative to $S$, and each such string can also be recognized by $M$ relative to $S - \{s_n\}$ (since

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s_n has Kolmogorov complexity greater than n/2, no machine can generate s_n from a string with only O(\log n) bits so in M's computation on a tally string of length at most q(n), s_n is never generated as a query string. Since every tall string in T of length at most q(n) can be recognized by M relative to S− \{s_n\}, the fact that S the strings in S whose length is of the form \(2^i\) are random allows us to conclude that every tally string in T of length at most q(n) can be generated in polynomial time by using c + O(\log n) bits, where c is a constant (that accounts for the short strings in S).

Recall that S \(\leq^p_T \) T. Since S \(\leq^p_T \) T and T \(\leq^p_T \) S, the argument in the last paragraph shows that S is self-reducible, contradicting the choice of S. (An alternative view is that for each n of the form \(2^i\), a machine generate in turn each string of length n and reduce each to T, thus discovering which of these strings is s_n. Thus, s_n can be determined by using O(\log n) bits, contradicting the choice of s_n as a string with Kolmogorov complexity greater than n/2.)

Recall that
\[P_T(TALLY) = P_H(TALLY) = P_H(SPARSE) = P_T(SPARSE) = P/poly.\]

We know that
\[E^p_T(TALLY) \neq E^p_T(SPARSE) \quad \text{and} \quad E^p_H(TALLY) \neq E^p_H(SPARSE).\]

In Theorem 3.5 below we separate \(E^p_H(TALLY)\) from \(E^p_T(TALLY)\) even though \(P_H(TALLY) = P_T(TALLY)\). The problem
\[E^p_H(SPARSE) = ? E^p_T(SPARSE)\]
remains open.

**Lemma 3.4:** If \(A \in E^p_H(TALLY)\), then \(P_T(A) = P_H(A)\).

**Proof:** This follows immediately from the fact that for any tally set T, \(P_T(T) = P_H(T)\). ■

**Theorem 3.5:** \(E^p_H(TALLY) \neq E^p_T(TALLY)\).

**Proof:** For any tally set \(T \subseteq \{0\}^*\), define \(A(T) = \{x_1 1 x_2 1 \ldots 1 x_t| \text{each } x_i \in \{0\}^*, \ t \geq 1, |x_1| < |x_2| < \ldots < |x_t|, \text{ and there exists } n \text{ such that } T^{\leq n} = \{x_1, x_2, \ldots, x_t\}\}. \) It is clear that \(A(T) \leq^p_T T\). Since \(A(T)\) is a set that
is linearly ordered, on input \(0^n\) a machine computing in polynomial time relative to \(A(T)\) can obtain the longest \(w \in A(T)\) such that \(|w|^2 \leq n^2\). From this \(w\) it is easy to determine whether \(0^n\) is in \(T\). Hence, \(T \leq^p_A(T)\). Thus, for every tally set \(T, A(T) \in E^p_T(T) \subseteq E^p_T(\text{TALLY})\). Thus, to prove that \(E^p_T(\text{TALLY}) \neq E^p_T(\text{TALLY})\), it is sufficient by Lemma 3.4 to show the existence of a tally set \(X\) such that \(X \not\leq^p_A(X)\).

We use the following notation: for any \(w \in \{0, 1\}^*\), if \(w = x_1 x_2 x_3 \ldots x_t\), \(t \geq 1\), each \(x_i \in \{0\}^*\), then tally-set \((w) = \{x_1, \ldots, x_t\}\).

[Recall that a set \(B\) is truth-table reducible to set \(C\), written \(B \leq^p_A C\), if there exist polynomial time computable functions \(f\) (the generator) and \(g\) (the evaluator) such that for all \(x\), \(f(x)\) is a list of strings, \(g(x)\) is a Boolean circuit with the number of input variables being equal to the number of strings in the list \(f(x)\), and \(x \in B\) if and only if \(g(x)\) evaluates to true on \(\langle \chi_C(y_1), \ldots, \chi_C(y_k) \rangle\) where \(f(x) = \langle y_1, \ldots, y_k \rangle\).]

The existence of a tally set \(X\) such that \(X \not\leq^p_A(X)\) can be shown by diagonalization. Since each function \(f\) and each function \(g\) making up reduction \(\langle f, g \rangle\) is polynomial time computable, we can assume an enumeration of \(\leq^p_A\)-reductions: \(\langle f_i, g_i \rangle | i \leq 1\) where for every \(i\), there is a machine that computes both \(f_i\) and \(g_i\) and has running time at most \(p_i(n) = n^i + i\).

The set \(X\) can be constructed by stages and only a sketch is presented.

Stage 0: \(X_0 := \emptyset\) and \(n_0 := 1\).

Stage \(i > 0\): Choose \(m > p_{i-1}(n_{i-1})\) and \(2m > p_i(2m) \geq t\). Let \(n_i := 2m\). At the beginning of this stage \(X_{i-1}\) is a tally set containing no string longer than \(n_{i-1}\).

If \(f_i(0^2 m) = \langle y_1, y_2, \ldots, y_t \rangle\), then choose \(S \subseteq \{0^m, 0^{m+1}, \ldots, 0^{2m-1}\}\) so that for every \(y_j, X_{i-1} \cup S \neq \text{tally-set}(y_j) - \{0^{2m}\}\) [the choice of \(m\) such that \(2m > p_i(2m) \geq t\) guarantees that such an \(S\) exists]. For such a set \(S\), observe that for every \(y_j, y_j \notin A(X_{i-1} \cup S)\) and \(y_j \notin A(X_{i-1} \cup S \cup \{0^{2m}\})\). Hence,

\[g_i(0^2 m, \chi_A(X_{i-1} \cup S)(f_i(0^2 m))) = g_i(0^2 m, \chi_A(X_{i-1} \cup S \cup \{0^{2m}\})(f_i(0^2 m))).\]

Finally, if

\[g_i(0^2 m, \chi_A(X_{i-1} \cup S)(f_i(0^2 m))) = 1,\]

then let \(X_i := X_{i-1} \cup S\); otherwise, let \(X_i := X_{i-1} \cup S \cup \{0^{2m}\}\).

This completes Stage \(i\).

Let \(X := \bigcup_{i \geq 0} X_i\).
From the construction, $0^V \in X$ if and only if $g_i(0^V, \chi_{A(\langle f_i, 0^V \rangle)}) = 0$. Thus, for every $i$, $\langle f_i, g_i \rangle$ does not witness $X \leq^P_{tt} A(X)$. Hence, $X \not\leq^P_{tt} A(X)$ as desired.

To obtain other separations from known results, we use the following simple fact.

**Lemma 3.6:** Let $C$ be a class of sets and let $\alpha$ and $\beta$ be two different reducibilities in $\{ \leq_T, \leq_{tt}, \leq_{bt}, \leq_k \text{ for each } k > 0, \leq_m \}$. If $P_\alpha(C) \neq P_\beta(C)$, then $E^p_\alpha(C) \neq E^p_\beta(C)$.

**Proof:** For any sets $A, B \subseteq \Sigma^*$, let $A \oplus B$ denote $\{ 0x, 1y \mid x \in A, y \in B \}$.

Suppose that there exists $A \in P_\alpha(C) - P_\beta(C)$. Since $A \in P_\alpha(C)$, there exists $C_0 \in C$ such that $A \leq^\alpha_\Sigma C_0$. Thus, $A \oplus C_0 \leq^\alpha_\Sigma C_0$ and $C_0 \leq^\beta_\Sigma A \oplus C_0$, so that $A \oplus C_0 \leq^\beta_\Sigma C_0$ and $A \oplus C_0 \in E^p_\alpha(C)$. Since $A \notin P_\beta(C)$, for every $C \in C$, $A \not\leq^\beta_\Sigma C$. If there exists $C_1 \in C$ such that $A \oplus C_0 \leq^\beta_\Sigma C_1$, then $A \oplus C_0 \leq^\beta_\Sigma C_1$ and, hence, $A \not\leq^\beta_\Sigma C_1$, contradicting the fact that $A \notin P_\beta(C)$. Thus, $A \oplus C_0 \notin E^p_\beta(C)$. Hence, $A \oplus C_0 \notin E^p_\beta(C)$. ■

Notice that $E^p_{bt}(\text{TALLY}) \neq E^p_{bt}(\text{SPARSE})$ by Theorem 3.2. Furthermore, $E^p_{bt}(\text{TALLY}) \neq E^p_{tt}(\text{TALLY})$ and $E^p_{bt}(\text{SPARSE}) \neq E^p_{tt}(\text{SPARSE})$ as in Theorem 3.7 (a) and (b) below.

**Theorem 3.7:**
(a) $E^p_{bt}(\text{TALLY}) \neq E^p_{tt}(\text{TALLY})$.
(b) $E^p_{bt}(\text{SPARSE}) \neq E^p_{tt}(\text{SPARSE})$.
(c) For every $k$, $E^p_{k-\text{tt}}(\text{SPARSE}) \neq E^p_{(k+1)-\text{tt}}(\text{SPARSE})$.
(d) $E^p_{m}(\text{SPARSE}) \neq E^p_{k-\text{tt}}(\text{SPARSE})$.
(e) $E^p_{ctt}(\text{SPARSE}) \neq E^p_{tt}(\text{SPARSE})$.
(f) $E^p_{ctt}(\text{SPARSE}) \neq E^p_{tt}(\text{SPARSE})$.
(g) $E^p_{ctt}(\text{SPARSE}) \neq E^p_{bt}(\text{SPARSE})$.
(h) $E^p_{tt}(\text{SPARSE}) \neq E^p_{bt}(\text{SPARSE})$.
(i) $E^p_{ctt}(\text{TALLY}) \neq E^p_{tt}(\text{TALLY})$.
(j) $E^p_{ctt}(\text{TALLY}) \neq E^p_{tt}(\text{TALLY})$.
(k) $E^p_{ctt}(\text{TALLY}) \neq E^p_{bt}(\text{TALLY})$.
(l) $E^p_{tt}(\text{TALLY}) \neq E^p_{tt}(\text{TALLY})$.
(m) $E^p_{tt}(\text{TALLY}) \neq E^p_{tt}(\text{TALLY})$.

The proof of Theorem 3.7 follows easily from Lemma 3.6 and by results of Book and Ko [5] and of Ko [9] on the separation of reduction classes.

The most interesting open questions are those stated in Section 1; they have been studied by Allender and Watanabe [2]. Other questions remain...
open; for example, the question of whether $E_{T}^{P}(\text{SPARSE})$ is properly included in $P/poly$.

*Note added in proof:* There are some recent results concerning the open problems. These results are reported in the following:


**REFERENCES**