

Deuxième Année du Cycle Supérieur (2CS)
Combinatorial Optimisation Assignment Report

The Traveling Salesman Problem Exact and Approximate Algorithms

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Introduction

The traveling salesman problem (which will be denoted by **TSP** for brevity's sake) is a classic problem in computer science. It is a typical example of a combinatorial optimization problem, that is, an optimization problem with a *discrete* solution space.

In its simplest form, the **TSP** asks the following question: “A salesman wants to take *the best¹ possible itinerary* between a set of cities, every city must be visited exactly once, and the salesman must start and finish at the same city. How can he find this itinerary?”

It is not difficult to see the practical use of solving the **TSP**. In fact, many important problems like vehicle routing, scheduling, array clustering [4], and circuit design [3] can be *expressed²* as **TSP** instances.

Furthermore, the **TSP** is of particular theoretical interest to complexity theory researchers, as its decision variant is a member of a very important family of decision problems called **NP**-complete problems.

In this document, we will introduce the **TSP**, investigate some of its properties and applications, and propose a few algorithms for solving it.

¹Usually “best” means shortest.

²Formally speaking, these problems can be *reduced* to **TSP**.

Symbols and Notation

Chapter 1

General Concepts

1.1 Computational Complexity

Throughout this document, we will often find ourselves in need of a method to objectively measure the difficulty of a problem or the efficiency of an algorithm. Fortunately, there exists an entire branch of theoretical computer science that addresses these very questions: *the theory of computational complexity*.

The theory of computational complexity formalizes the intuitive concept of the *difficulty* of a problem. Quite reasonably, this discipline relies on the premise that a problem is as difficult as it is to perform its most efficient solution, or, to use technical terms, to *execute the most efficient algorithm* that solves the problem.

1.1.1 Decision Problems and Complexity Classes

For historical (and technical) reasons, most of the work in this branch has been done around a special type of problem called *decision problems*. A decision problem is a problem that has a binary answer, that is, given an instance (or an input) of the problem, we compute an answer (or output) that is an element of some preknown set with cardinality 2. The sets $\{0, 1\}$, $\{\text{false}, \text{true}\}$, and $\{\text{no}, \text{yes}\}$, are common examples of such a set, the former of which will be used in the rest of this discussion.

A decision problem can therefore be defined as the problem of evaluating some

computable¹ function $f : S \rightarrow \{0, 1\}$ where S is some set of inputs.

These functions can be mapped to decidable subsets of S by associating every such a set P with its characteristic function $\mathbb{1}_P$. This correspondence is what motivates Definition 1 of decidable problems.

Definition 1 (Decision Problem).

A decision problem is a pair $X = \langle S, P \rangle$ where S is a countable set called the *instance space* and $P \subset S$ is called the set of *positive instances*. We say the problem X is decidable iff P is decidable (i.e. if $\mathbb{1}_P$ is computable).

To better understand Definition 1, we consider Examples 1.1, and 1.2, the latter of which is of particular historical significance in the context of complexity theory.

Example 1.1 (A Number of Decision Problems).

- **PRIME** is the problem $\langle \mathbb{N}, P \rangle$ where P is the set of all prime numbers.
- **HAM**, or the *hamiltonicity problem* is the problem $\langle S, P \rangle$ where S is the set of all undirected graphs and P is the set of all hamiltonian graphs.
- **CLIQUE**. An instance of this problem is a pair $\langle G, k \rangle$ where G is a graph and $k \in \mathbb{N}$. Such an instance is positive iff G has a clique of size k .
- **PAIR** or the parity problem is the problem $\langle \{0, 1\}^*, P \rangle$ where

$$P = \{w \in \{0, 1\}^* \mid |w|_1 \equiv 0 \pmod{2}\}$$

Example 1.2 (SAT).

The *satisfiability problem of boolean logic*, or **SAT**, is the problem $\langle S, P \rangle$ where S is the set of all formulae of boolean logic and P is the set of all *satisfiable* formulae. A formula $\varphi \in S$ is satisfiable iff it has a satisfying assignment or a *model*, that is, if its negation is not a tautology. For instance, the formula φ below is satisfiable (by the assignment $x_1 = 0, x_2 = 0, x_3 = 1$ for example) while ψ is not.

$$\begin{aligned}\varphi &= ((x_1 \vee \neg x_2) \wedge x_3) \vee (x_2 \wedge \neg x_3) \\ \psi &= \neg(x_1 \vee \neg x_2) \wedge \neg(x_2 \vee \neg x_1)\end{aligned}$$

¹The model of computation is not important when defining decision problems, but it becomes so when discussing their complexity. The turing machine is the model we will use throughout this work.

The complexity of a decision problem is given by two pieces of information, the first being its time complexity (intuitively, this is the runtime of the *fastest* turing machine deciding the problem), and the second its space complexity (the *minimal* number of distinct visited cells on the tape of turing machine deciding the problem). The rigorous definitions of these quantities are given in Appendix A. We will use the notions of *algorithms*, *runtime*, and *memory usage* as intuitive analogues of Turing machines, runtime and space complexity respectively.

As is common when discussing complexity, we will sort problems in a hierarchy of *complexity classes*. These complexity classes are based on the asymptotic behavior of the time and space complexities instead of the exact runtime or memory usage of a particular algorithm solving a particular problem. A few important complexity classes are given by Definition 2 [1].

Definition 2 (Most Important Complexity Classes).

Let $f : \mathbb{N} \rightarrow \mathbb{R}_+$ be a function. We define the following classes of problems:

- $\text{TIME}(f(n))$ is the set of problems X for which there exists a deterministic Turing machine \mathcal{M} that decides X such that $t_{\mathcal{M}}(n) = O(f(n))$.
- $\text{NTIME}(f(n))$ is the set of problems X for which there exists a nondeterministic Turing machine \mathcal{M} that decides X such that $t_{\mathcal{M}}(n) = O(f(n))$.

From these two, we can define the following classes:

$$\begin{aligned} \text{P} &= \bigcup_{k \in \mathbb{N}} \text{TIME}(n^k) & \text{NP} &= \bigcup_{k \in \mathbb{N}} \text{NTIME}(n^k) \\ \text{EXPTIME} &= \bigcup_{k \in \mathbb{N}} \text{TIME}(2^{n^k}) & \text{NEXPTIME} &= \bigcup_{k \in \mathbb{N}} \text{NTIME}(2^{n^k}) \end{aligned}$$

The list given by Definition 2 is of course far from complete. In fact, a very easy way to extend it is to add for every class C its *dual* class $\text{co-}C := \{\bar{X} \mid X \in C\}$ of complements of problems in C . It is however largely sufficient for the purposes of this investigation. In fact, we will mostly be dealing with the classes P and NP exclusively.

It is quite straightforward to verify the following inclusions between the classes defined thus far:

$$\begin{aligned} \text{P} &\subset \text{NP} \subset \text{EXPTIME} \subset \text{NEXPTIME} \\ \text{P} &\subset \text{co-NP} \subset \text{EXPTIME} \subset \text{co-NEXPTIME} \end{aligned}$$

However, it is exceedingly difficult to prove strict inclusion for most pairs. As a matter of fact, the problem of determining whether $P = NP$ is an open one, as well as that of finding the intersection between C and $\text{co-}C$ for $C \in \{NP, \text{NEXPTIME}\}$. The Venn diagram of Figure 1.1 shows the known inclusions (opting for strictness for unknown pairs).

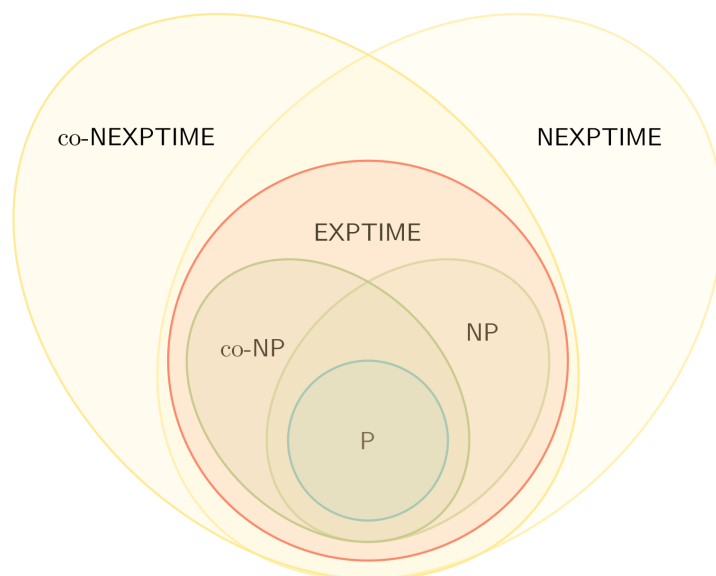


Figure 1.1: A Venn diagram of the classes defined so far.

1.1.2 Reducibility and Completeness

As we have seen in the previous section, it is currently unknown whether $P = NP$. As far as we know, it could be as easy to decide a problem with a deterministic machine as it is to decide it with a nondeterministic one. We don't know if problems in NP are *strictly harder* than those in P (even though we suspect them to be). This is hardly surprising given the vastness of NP . In order to prove $P = NP$, one must find a polynomial time algorithm for *every* problem in NP . And in order to prove that $P \neq NP$, one must prove that for at least one problem in NP , *all algorithms* are not polynomial time (this is admittedly an easier task than proving equality).

The point above is valid for all pairs of classes for which we have one inclusion but are uncertain of the other (NP and $EXPTIME$ for example). The notion of

completeness simplifies the task of proving inequality of two classes by reducing it to the task of proving that *one particular* problem (intuitively thought of as the hardest problem) in the supposedly bigger class is in deed a member of the smaller one. These *most difficult problems* are what the concept of completeness aims to formalize. In order to define them, we must first address the question of how to compare the difficulty of two problems, which is exactly what the relation of *reducibility* introduced in Definition 3 allows us to do.

Definition 3 (Polynomial Reduction).

Let $X = \langle S_1, P_1 \rangle$ and $Y = \langle S_2, P_2 \rangle$ be two decision problems. A *polynomial reduction* of X to Y is a function $f : S_1 \rightarrow S_2$ computable by a deterministic Turing machine in polynomial time such that:

$$\forall i \in S_1 \ i \in P_1 \iff f(i) \in P_2$$

If such a reduction exists, X is said to be *polynomially reducible* to Y , which we denote by $X \preceq_P Y$.

One can show that \preceq_P is both transitive, and reflexive (a preorder on decision problems) without too much difficulty. Furthermore, the relation \preceq_P can be shown to satisfy Proposition 1.1 which will become very important once NP-completeness is defined.

Proposition 1.1.

Let X and Y be two decision problems. If $X \preceq_P Y$ and $Y \in P$, then $X \in P$.

Proof.

The composition of the Turing machine that computes the polynomial reduction of X to Y and the one that decides Y in polynomial time is a polynomial time Turing machine that decides X . \square

We now define NP-completeness.

Definition 4 (NP-hardness, NP-completeness).

A problem X is NP-hard if and only if :

$$\forall Y \in \text{NP} \ Y \preceq_P X$$

A problem X is NP-complete if and only if X is NP-hard and $X \in \text{NP}$.

Under our intuitive interpretation of \preceq_P , an NP-hard problem is a problem that is at least as difficult as any problem in NP. An NP-complete one is then the hardest problem in NP. It stands to reason then that if an NP-hard problem X is in P, we would have $NP \subset P$ and hence $P = NP$. This is in deed the case as affirmed by Corollary 1.2.

Corollary 1.2. *If $P \cap NP\text{-hard} \neq \emptyset$, then $P = NP$*

Proof.

Let X be an NP-hard problem that is also in P, and let $Y \in NP$. By definition of NP-hardness, $Y \preceq_P X$, and by Proposition 1.1 and the fact that $X \in P$, $Y \in P$. Therefore, $NP \subset P$, and hence $P = NP$. \square

The same intuitive analysis that led us to Corollary 1.2 suggests that if a problem is harder² than an NP-hard problem, then it must be NP-hard as well. This is once again correct as affirmed by Proposition 1.3.

Proposition 1.3.

Let X and Y be two decision problems. If X is NP-hard and $X \preceq_P Y$, then Y is NP-hard.

Proof.

Let X, Y, Z be decision problems such that X is NP-hard and $X \preceq_P Y$, and $Z \in NP$. By NP-hardness of X , $Z \preceq_P X$, and by transitivity of \preceq_P we have $Z \preceq_P Y$. Y is then NP-hard. \square

Although the idea of NP-completeness is promising, it is not immediately obvious how it makes approaching P vs NP easier. It would seem that proving a problem is NP-hard is as difficult as solving P vs NP. This is once more due to the unfathomable vastness of NP. Fortunately, this initial impression proved wrong. We know of many NP-hard and NP-complete problems. The first problem to be proven NP-complete is the problem SAT we invoked in Example 1.2. This has been done by Stephen Cook and Leonid Levin who proved Theorem 1.4 in 1971. A proof of this theorem is given in [2].

Theorem 1.4 (Cook Levin).

SAT is NP-complete.

²In the \preceq_P sense.

Now that we have one problem that we know is **NP**-complete, it becomes significantly easier to prove that a given problem is **NP**-hard. Using Proposition 1.3, we can show any problem is **NP**-hard by reducing a known **NP**-hard problem to it. This is what we will do in the following example by proving 3-SAT is **NP**-complete.

Example 1.3.

An instance of 3-SAT is a conjunction of clauses with at most 3 literals each and is positive iff it is satisfiable. To show that 3-SAT is **NP**-complete, we will reduce SAT to it. It is technically necessary to show that 3-SAT \in **NP** too, but this is easy given that it is a subproblem of SAT which is already known to be in **NP**. We will therefore focus on proving it **NP**-hard.

To reduce SAT to 3-SAT, we must find for every formula φ of boolean logic a formula φ' in 3CNF, with size polynomial in that of φ , such that φ' is satisfiable iff φ is satisfiable. We start by considering the *syntax tree* of φ , for example, the syntax tree of the formula

$$\varphi = ((x_1 \vee \neg x_2) \wedge x_3) \vee (x_2 \wedge \neg x_3)$$

is shown in Figure 1.2. We label each internal node of this tree with a new variable (the leaves are associated with the variables of φ), this produces the tree of Figure 1.3.

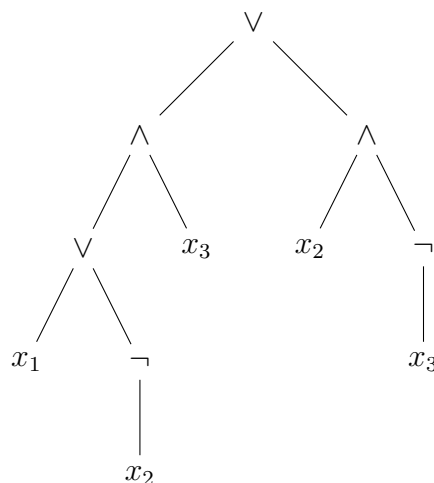


Figure 1.2: Syntax tree of the formula φ

Each new variable x_i is then associated with an equivalence e_i according to the following rules:

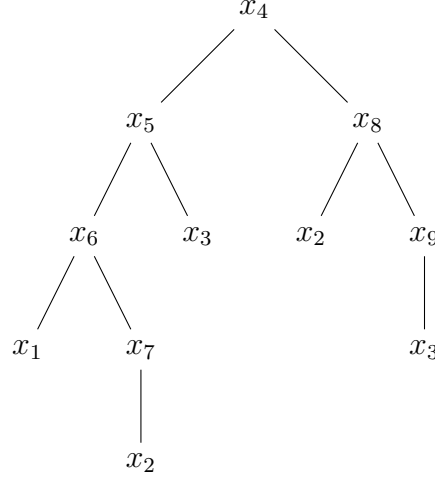


Figure 1.3: Syntax tree of Figure 1.2 labeled with new variables.

- If the node labeled by the variable x_i is \vee , we associate it with the formula $x_i \leftrightarrow x_j \vee x_k$ where x_j and x_k are the children of x_i .
- If it is labeled with \wedge , we associate it with the formula $x_i \leftrightarrow x_j \wedge x_k$.
- If it is labeled with \neg , we associate with the formula $x_i \leftrightarrow \neg x_j$.

Finally, by taking the conjunction³ $\bigwedge_{i>n} e_i$ of the equivalences thus obtained, and replacing equivalences with their clausal forms given by the following tautologies:

$$\begin{aligned}
 x \leftrightarrow \neg y &\equiv (x \vee \neg y) \wedge (\neg x \vee y) \\
 x \leftrightarrow y \wedge z &\equiv (x \vee \neg y \vee \neg z) \wedge (\neg x \wedge y) \wedge (\neg x \wedge z) \\
 x \leftrightarrow y \vee z &\equiv (\neg x \vee y \vee z) \wedge (x \wedge \neg y) \wedge (x \wedge \neg z)
 \end{aligned}$$

We get a formula φ' in 3CNF, which is logically equivalent to φ (and is a fortiori satisfiable iff φ is satisfiable). Furthermore, since the size of the syntax tree is proportional to the size of φ , so is the size of φ' (i.e. this reduction is indeed polynomial). It then follows that $\text{SAT} \preceq_P \text{3-SAT}$, and by Proposition 1.3, that 3-SAT is NP-complete.

Using similar techniques, countless problems have been shown to be NP-complete. Examples include **CLIQUE** (the problem of deciding whether a graph has a k -clique

³ n being the number of variables in φ

for $k \in \mathbb{N}$), VERTEX-COVER (the problem of deciding whether a graph has a covering set of vertices with cardinality k for $k \in \mathbb{N}$), and crucially for our study, HAM (mentioned in Example 1.1).

1.1.3 Optimization Problems and Their Classification

Despite their ubiquity, decision problems are not always powerful enough to model a given situation. Combinatorial optimization extends them with a much more general class of problems. Instead of finding a binary answer, a combinatorial optimization problem (or simply, an optimization problem) seeks an *optimal* solution in a discrete *space of options*. The optimality of a solution is judged with respect to an *objective function*.

Definition 5 (Optimization Problem).

An optimization problem is a quintuplet $A = \langle X, S, F, \mu, g \rangle$ where

- (i) X is a set of *instances*,
- (ii) S is a set of *solutions*,
- (iii) $F : X \rightarrow \mathcal{P}(S)$, for $x \in X$, $F(x)$ is called the set of feasible solutions of instance x ,
- (iv) $\mu : \{(x, s) \mid x \in X \wedge s \in F(x)\} \rightarrow \mathbb{R}_+$, $\mu(x, s)$ is called the *measure* of the solution s to instance x ,
- (v) $g \in \{\min, \max\}$ is the *goal* of the problem.

For an instance x , we define the *objective function* $\mu_x : F(x) \rightarrow \mathbb{R}_+$, $s \mapsto \mu(x, s)$, the *optimum* $\text{opt}(x) := g \{\mu(x, s) \mid s \in F(x)\}$, and the set of *optimal solutions* $\mu_x^{-1}(\{\text{opt}(x)\})$. Such an instance is said to be *solvable* if $F(x) \neq \emptyset$.

By *solving* A , we mean finding a computable function $f : X \rightarrow S$ (also called a solver) such that for every solvable $x \in X$, $f(x) \in F(x)$. If f has the additional property that for every solvable $x \in X$, $f(x)$ is an optimal solution to x , we say that f is an *exact solver* of A .

Chapter 2

Problem Statement

2.1 History

The first use of the term 'traveling salesman problem' in mathematical circles may have been in 1931-32, as we shall explain below. But in 1832, a book was printed in Germany entitled *Der Handlungsreisende, wie er sein soll und was er zu thun hat, um Aufträge zu erhalten und eines glücklichen Erfolgs in seinen Geschäften gewiss zu sein. Von einem alten Commis-Voyageur* ("The Traveling Salesman, how he should be and what he should do to get Commissions and to be Successful in his Business. By a veteran Traveling Salesman").

Although devoted for the most part to other issues, the book reaches the essence of the TSP in its last chapter: 'By a proper choice and scheduling of the tour, one can often gain so much time that we have to make some suggestions.... The most important aspect is to cover as many locations as possible without visiting a location twice ...' [Voigt, 1831; MiMer-Merbach, 1983].

Chapter 3

Branch and Bound

3.1 Motivation

The direct method as we have seen is, despite the simplicity of its implementation, unrealistically slow for even very small instances.

Faster exact algorithms exist, but none of them is polynomial since TSP is NP-complete. In fact, under the assumption $P \neq NP$, no polynomial solution exists.

Branch and Bound is one such algorithm that we will dedicate the rest of the chapter to.

3.2 The idea of Branch and Bound

The idea of Branch and Bound is to eliminate certain branches from the search space to decrease runtime.

This is done by computing a *lower bound* and *upper bound* for every branch, and then pruning branches that are guaranteed to be worse than the best known solution.

3.3 The implementation

Appendix A

Computability

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